Abstract

Atomicity or strong consistency is one of the fundamental, most intuitive, and hardest to provide primitives in distributed shared memory emulations. To ensure survivability, scalability, and availability of a storage service in the presence of failures, traditional approaches for atomic memory emulation, in message passing environments, replicate the objects across multiple servers. Compared to replication based algorithms, erasure code-based atomic memory algorithms has much lower storage and communication costs, but usually, they are harder to design. The difficulty of designing atomic memory algorithms further grows, when the set of servers may be changed to ensure survivability of the service over software and hardware upgrades (scale-up or scale-down), while avoiding service interruptions. Atomic memory algorithms for performing server reconfiguration, in the replicated systems, are very few, complex, and are still part of an active area of research; reconfigurations of erasure-code based algorithms are non-existent.

In this work, we present ARES, an algorithmic framework that allows reconfiguration of the underlying servers, and is particularly suitable for erasure-code based algorithms emulating atomic objects. ARES introduces new configurations while keeping the service available. To use with ARES we also propose a new, and to our knowledge, the first two-round erasure-code based algorithm TREAS, for emulating multi-writer, multi-reader (MWMR) atomic objects in asynchronous, message-passing environments, with near-optimal communication and storage costs. Our algorithms can tolerate crash failures of any client and some fraction of servers, and yet, guarantee safety and liveness property. Moreover, by bringing together the advantages of ARES and TREAS, we propose an optimized algorithm where new configurations can be installed without the objects values passing through the reconfiguration clients.
1 Introduction

With the rapid increase of computing power on portable devices, such as, smartphone, laptops, tablets, etc., and the near ubiquitous availability of Internet connectivity, our day to day lives are becoming increasingly dependent on Internet-based applications. Most of these applications, rely on large volumes of data from a wide range of sources, and their performance improves with the easy accessibility of the data. Today, data is gathered at an even faster pace from numerous sources of interconnected devices around the world. In order to keep abreast with this veritable tsunami of data, researchers, in both industry and academia, are hurrying to invent new ways to increase the capacity of durable, large-scale distributed storage systems, and the efficient ingestion and retrieval of data. Currently, most of the data is stored in cloud-based storages, offered by major providers like Amazon, Dropbox, Google, etc. To store large amounts of data in an affordable manner, cloud vendors deploy hundreds to thousands of commodity machines, networked together to act as a single giant storage system. Component failures of commodity devices, and network delays are the norm, therefore, ensuring consistent data-access and availability at the same time is challenging. Vendors often solve availability by generating replicating data across multiple servers, but this creates the headache of keeping those copies consistent, especially when they can be updated concurrently by different operations.

Atomic Distributed Storage. To solve this problem, researchers developed a series of consistency semantics, that formally specify the behavior of concurrent operations. Atomicity or strong consistency is the strongest and most intuitive consistency semantic which provides the illusion that a data object is accessed sequentially, even when operations may access different copies of that data object concurrently. In addition, atomic objects are composable \[\text{17, 24}\], enabling the creation of large shared memory systems from individual atomic data objects. Such large-scale shared memory services make application development much simpler. Finally, the strongest and most desirable form of availability or liveness of atomic data access is wait-freedom \[\text{24}\] where operations complete irrespective of the speed of the clients.

Replication-based Atomic Storage. A long stream of works used replication of data across multiple servers to implement atomic read/write storage \[\text{3, 4, 9, 10, 11, 13, 14, 26}\]. Popular replication-based algorithms appear in the work by Attiya, Bar-Noy and Dolev \[4\] (we refer to this as the ABD algorithm) and also in the work by Fan and Lynch \[10\] (which is referred to as the LDR algorithm). Replication based strategies, however, incur high storage and communication costs; for example, to store 1,000,000 objects each of size 1MB, a total size of 1 TB across a 3 server system, the ABD algorithm replicates the objects in all the 3 servers, which blows up the worst-case storage cost to 3 TB. Additionally, every write or read operation may need to transmit at most 3 MB of data, incurring high communication cost.

Erasure Code-based Atomic Storage. To avoid the high storage and communication costs stemmed from the use of replication, erasure codes provide an alternative way to emulate fault-tolerant shared atomic storage. In comparison to replication, algorithms based on erasure codes significantly reduce both the storage and communication costs of the implementation. An \([n, k]\) erasure code splits the value \(v\) of size 1 unit into \(k\) elements, each of size \(\frac{1}{k}\) units, creates \(n\) coded elements, and stores one coded element per server. The size of each coded element is also \(\frac{1}{k}\)
units, and thus the total storage cost across the $n$ servers is $\frac{2}{k}$ units. For example, if we use an $[n = 3, k = 2]$ MDS code, the storage cost is simply 1.5 TB, which is 2 times lower than the storage in the case of ABD. A similar reduction in bandwidth per operation is possible in many erasure code-based algorithms for implementing atomicity. A class of erasure codes known as Maximum Distance Separable (MDS) codes have the property that value $v$ can be reconstructed from any $k$ out of these $n$ coded elements. Motivated by these approaches, recently, several ensure code based algorithms for implementing strong consistency on message-passing models have been proposed in theory \cite{5, 6, 8, 20, 22}, and in practice \cite{7, 29}. However, the leap from replication-based to erasure code-based algorithms for strong consistency comes with the additional burden of synchronizing the access of multiple pieces of coded-elements from the same version of the data object. This naturally leads to relatively complex algorithms.

Reconfigurable Atomic Storage. Although replication and erasure-codes may help the system survive the failure of a subset of servers, they do not suffice to ensure the liveness of the service in a longer period where a larger number of servers may fail. Reconfiguration is the process that allows addition or removal of servers from a live system, without affecting its normal operation. However, forming reconfiguration of a system, without service interruption, is very challenging and not well-understood even in replicated systems, and is still an active area of research. RAMBO \cite{25} and DynaStore \cite{1} are two of the handful of algorithms \cite{12, 15, 19, 27} that allows reconfiguration on live systems. Recently, the authors in \cite{28} presented a general algorithmic framework for consensus-free reconfiguration algorithms.

So far, all reconfiguration approaches, implicitly or explicitly, assume a replication-based system in each configuration. Therefore, none of the above methods can reconfigure from or to a configuration where erasure codes are used. In particular, erasure code based atomic memory algorithms requires a fixed set of servers in a configuration, therefore, moving from an existing coding scheme would entail deploying a new set of servers, or change the set of servers in an existing configuration. In RAMBO \cite{25}, messages originating from a read or write operation are sent as part of an ongoing gossiping protocol, and this makes counting the communication cost of each operation obscure. In both RAMBO \cite{25} and DynaStore \cite{1}, the clients and servers are combined, and therefore, not immediately suitable for settings where clients and servers are separate processes, a commonly studied architecture, both in theory \cite{24} and in practice \cite{23}. In DynaStore \cite{1} and \cite{28}, an active new configuration, generated by the underlying SpSn algorithm, may often consist of a set of servers proposed during configuration proposed by multiple clients. They assume that as long as a majority of the servers in the active configurations are non-faulty the service can guarantee liveness. In erasure code based algorithms additional assumptions are required. For example, if some client proposes a configuration with certain MDS $[n, k]$ code then if we have more than $n$ servers in the configuration increases the storage cost, while having fewer than $n$ servers will not permit using the proposed code parameters. Moreover, in the algorithms in \cite{1} and \cite{28}, the configurations proposed by the clients are incremental changes, e.g., $\{-s_1, -s_2, +s_3\}$, where $-$ is a suggestion to remove a server and $+$ a suggestion to add it. Therefore, some additional mechanism is necessary for the reconfiguration client to know the total number of existing active servers before proposing its change as an attempt to generate a configuration of a desired size.

Reconfigurations are usually desirable to system administrators \cite{2}, usually during system maintenance. Therefore, during the migration of the objects, from one configuration to the next, it is highly likely that all stored objects are moved to the newer configuration almost at the same
time. In the above algorithms, data transfer is done, by using the client as a conduit, creating a possible bottleneck. Although, setting proxy servers for reconfiguration is an ad hoc solution it suffers from a the same bottleneck. In such situations, it is more reasonable for the data objects to migrate directly from one set of servers to another.

**Our Contributions.** In this work, we present ARES, an algorithm that allows reconfiguration of the servers that emulates an atomic memory, specifically suitable for atomic memory service that uses erasure codes, while keeping the service available at all times. In order to keep ARES general, so as to allow adaptation of already known atomic memory algorithms to the configurations, we introduced a new set of data primitives, that (i) provides a modular framework to describe atomic read/write storage implementations, and (ii) enhances the ease for reasoning about algorithm’s correct implementation of atomic memory. Using these primitives, we are able to prove the safety property (atomic memory) of an execution of ARES that involves ongoing reconfiguration operations. We also present TREAS, the first two-round erasure code-based MWMR algorithm, with cost-effective storage and communication costs, for emulating shared atomic memory under message-passing environment in the presence of crash-failure of processes. We prove safety and liveness conditions for TREAS. Then we describe a new algorithm ARES-TREAS, where we use a modified version of TREAS as the underlying atomic memory algorithm in every configuration and modify ARES, so that data from one configuration is transferred to another directly, thereby, avoiding the reconfiguration client as the possible bottleneck.

**Document Structure.** The remainder of the manuscript consists of the following sections. In Section 2 we present the model assumptions for our setting. In Section 3 we present our two-round erasure code-based algorithm for emulating shared atomic storage under the message-passing model for MWMR setting. In Section 4 we present our ARES framework for emulating shared atomic memory with erasure-codes where the system can undergo reconfiguration, while it is live. Sub-section 4.4 provides latency analysis read, write and reconfiguration operations. In Section 5 we describe ARES-TREAS algorithm. Finally, in Section 6 we conclude our work.

## 2 Model and Definitions

A shared atomic storage can be emulated by composing individual atomic objects. Therefore, we aim to implement only one atomic read/write memory object on a set of servers. Each data object takes a value from a set \( \mathcal{V} \). We assume a system consisting of four distinct sets of processes: a set \( \mathcal{W} \) of writers, a set \( \mathcal{R} \) of readers, a set \( \mathcal{G} \) of reconfiguration clients, and a set \( \mathcal{S} \) of servers. Let \( \mathcal{I} = \mathcal{W} \cup \mathcal{R} \cup \mathcal{G} \) be the set of clients. Servers host data elements (replicas or encoded data fragments). Each writer is allowed to modify the value of a shared object, and each reader is allowed to obtain the value of that object. Reconfiguration clients attempt to modify the set of servers in the system in order to mask transient errors and to ensure the longevity of the service. Processes communicate via messages through asynchronous, reliable channels. ARES allows the set of server host to be modified during the course of an execution for masking transient or permanent failures of servers and preserve the longevity of the service.

**Configuration** A configuration, identified by a unique identifier \( c \), is a data type that describes explicitly: (i) the set identifiers of the set of servers that are a part of it, denote as \( c.Servers \); (ii)
the set of quorums that are defined on $c.Servers$; $(iii)$ an underlying algorithm that implements atomic memory in $c.Servers$, including related parameters; and $(iv)$ a consensus instance, $c.Con$, with the values from $C$, the set of all unique configuration identifiers, implemented on top of the servers in $c.Servers$. We refer to a server $s \in c.Servers$ as a member of configuration $c$.

**Liveness and Safety Conditions.** The algorithms we present in this paper satisfy wait-free termination (Liveness) and atomicity (Safety). Wait-free termination \cite{16} requires that any process terminates in a finite number of steps, independent of the progress of any other process in the system.

An implementation $A$ of a read/write object satisfies the atomicity property if the following holds \cite{24}. Let the set $\Pi$ contain all complete operations in any well-formed execution of $A$. Then for operations in $\Pi$ there exists an irreflexive partial ordering $\prec$ satisfying the following:

A1. For any operations $\pi_1$ and $\pi_2$ in $\Pi$, if $\pi_1 \rightarrow \pi_2$, then it cannot be the case that $\pi_2 \prec \pi_1$.

A2. If $\pi \in \Pi$ is a write operation and $\pi' \in \Pi$ is any operation, then either $\pi \prec \pi'$ or $\pi' \prec \pi$.

A3. The value returned by a read operation is the value written by the last preceding write operation according to $\prec$ (or the initial value if there is no such write).

**Storage and Communication Costs.** We are interested in the complexity of each read and write operation. The complexity of each operation $\pi$ invoked by a process $p$, is measured with respect to three metrics, during the interval between the invocation and the response of $\pi$: (i) communication round-trips, accounting the number of messages sent during $\pi$, (ii) storage efficiency (storage cost), accounting the maximum storage requirements for any single object at the servers during $\pi$, and (iii) message bit complexity (communication cost) which measures the length of the messages used during $\pi$.

We define the total storage cost as the size of the data stored across all servers, at any point during the execution of the algorithm. The communication cost associated with a read or write operation is the size of the total data that gets transmitted in the messages sent as part of the operation. We assume that metadata, such as version number, process ID, etc. used by various operations is of negligible size, and is hence ignored in the calculation of storage and communication cost. Further, we normalize both the costs with respect to the size of the value $v$; in other words, we compute the costs under the assumption that $v$ has size 1 unit.

**Background on Erasure coding.** In TREAS, we use an $[n, k]$ linear MDS code \cite{18} over a finite field $\mathbb{F}_q$ to encode and store the value $v$ among the $n$ servers. An $[n, k]$ MDS code has the property that any $k$ out of the $n$ coded elements can be used to recover (decode) the value $v$. For encoding, $v$ is divided* into $k$ elements $v_1, v_2, \ldots, v_k$ with each element having size $\frac{1}{k}$ (assuming size of $v$ is 1). The encoder takes the $k$ elements as input and produces $n$ coded elements $c_1, c_2, \ldots, c_n$ as output, i.e., $[c_1, \ldots, c_n] = \Phi([v_1, \ldots, v_k])$, where $\Phi$ denotes the encoder. For ease of notation, we simply write $\Phi(v)$ to mean $[c_1, \ldots, c_n]$. The vector $[c_1, \ldots, c_n]$ is referred to as the codeword corresponding

*In practice $v$ is a file, which is divided into many stripes based on the choice of the code, various stripes are individually encoded and stacked against each other. We omit details of represent-ability of $v$ by a sequence of symbols of $\mathbb{F}_q$, and the mechanism of data striping, since these are fairly standard in the coding theory literature.
to the value $v$. Each coded element $c_i$ also has size $\frac{1}{k}$. In our scheme we store one coded element per server. We use $\Phi_i$ to denote the projection of $\Phi$ on to the $i$th output component, i.e., $c_i = \Phi_i(v)$. Without loss of generality, we associate the coded element $c_i$ with server $i$, $1 \leq i \leq n$.

Tags. A tag $\tau$ is defined as a pair $(z, w)$, where $z \in \mathbb{N}$ and $w \in \mathcal{W}$, an ID of a writer. Let $\mathcal{T}$ be the set of all tags. Notice that tags could be defined in any totally ordered domain and given that this domain is countably infinite, then there can be a direct mapping to the domain we assume. For any $\tau_1, \tau_2 \in \mathcal{T}$ we define $\tau_2 > \tau_1$ if (i) $\tau_2.z > \tau_1.z$ or (ii) $\tau_2.z = \tau_1.z$ and $\tau_2.w > \tau_1.w$.

2.1 The Data-Access Primitives

In the next section, we present the TREAS algorithm using three data access primitives (DAP) which can be associated with a configuration $c$. These data access primitives in the context of $c$ are: (i) $c$.put-data($\langle \tau, v \rangle$), (ii) $c$.get-data(), and (iii) $c$.get-tag(). Assuming a set of totally ordered timestamps $\mathcal{T}$, a value domain of the distributed atomic object $\mathcal{V}$, and a set of configuration identifiers $\mathcal{C}$, the three primitives defined over a configuration $c \in \mathcal{C}$, tag $\tau \in \mathcal{T}$, and a value $v \in \mathcal{V}$ as follows:

**Definition 1** (Data Access Primitives). Given a configuration identifier $c \in \mathcal{C}$, any non-faulty client process $p$ may invoke the following data access primitives during an execution $\xi$, where $c$ is added to specify the configuration specific implementation of the primitives:

**D1.** $c$.get-tag(): returns a tag $\tau \in \mathcal{T}$

**D2.** $c$.get-data(): returns a tag-value pair $(\tau, v) \in \mathcal{T} \times \mathcal{V}$

**D3.** $c$.put-data($\langle \tau, v \rangle$): the tag-value pair $(\tau, v) \in \mathcal{T} \times \mathcal{V}$ as argument.

**Algorithm 1** Read and write operations of generic algorithm $A_1$

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<table>
<thead>
<tr>
<th>Operation</th>
<th>Read</th>
</tr>
</thead>
<tbody>
<tr>
<td></td>
<td>2: $(t, v) \leftarrow c$.get-data()</td>
</tr>
<tr>
<td></td>
<td>4: return $(t, v)$</td>
</tr>
<tr>
<td></td>
<td>end operation</td>
</tr>
</tbody>
</table>

6: operation write$(v)$

8: $t_w \leftarrow inc(t)$

10: end operation

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A number of known tag-based algorithms that implement atomic read/write objects (e.g., ABD, FAST – see Appendix A), can be expressed in terms of DAP. In particular, many algorithms can be transformed in an algorithmic template, say $A_1$, presented in Alg. 10. In brief, a read operation in $A_1$ performs $c.c$.get-data() to retrieve a tag-value pair, $\langle \tau, v \rangle$ form configuration $c$, and then it performs a $c$.put-data($\langle \tau, v \rangle$) to propagate that pair to configuration $c$. A write operation is similar to the read but before performing the put-data action it generates a new tag which associates with the value to be written. It can be shown (see Appendix A) that an algorithm described in the form $A_1$ satisfies atomic guarantees and liveness, if the DAP satisfy the following consistency properties:

**Definition 2** (DAP Consistency Properties). In an execution $\xi$ we say that a DAP operation in an execution $\xi$ is complete if both the invocation and the matching response step appear in $\xi$. If $\Pi$ is the set of complete DAP operations in execution $\xi$ then for any $\phi, \pi \in \Pi$:
C1 If $\phi$ is $c$.put-data($\langle \tau_\phi, v_\phi \rangle$), for $c \in C$, $\tau_1 \in T$ and $v_1 \in V$, and $\pi$ is $c$.get-tag() (or $c$.get-data()) that returns $\tau_\pi \in T$ (or $\langle \tau_\pi, v_\pi \rangle \in T \times V$) and $\phi$ completes before $\pi$ in $\xi$, then $\tau_\pi \geq \tau_\phi$.

C2 If $\phi$ is a $c$.get-data() that returns $\langle \tau_\pi, v_\pi \rangle \in T \times V$, then there exists $\pi$ such that $\pi$ is $c$.put-data($\langle \tau_\pi, v_\pi \rangle$) and $\phi$ did not complete before the invocation of $\pi$, and if no such $\pi$ exists in $\xi$, then $\langle \tau_\pi, v_\pi \rangle$ is equal to $(t_0, v_0)$.

Expressing an atomic algorithm in terms of the DAP primitives allows one to achieve a modular design for atomic object tag-based algorithms, serving multiple purposes. First, describing an algorithm according to templates $A_1$ (like TREAS in Section 3) allows one to prove that the algorithm is safe (atomic) by just showing that the appropriate DAP properties hold, and the algorithm is live if the implementation of each primitive is live. Second, the safety and liveness proofs for more complex algorithms (like ARE in Section 4) become easier as one may reason on the DAP properties that are satisfied by the primitives used, without involving the underlying implementation of those primitives. Last but not least, describing a reconfigurable algorithm using DAPs, provides the flexibility to use different implementation mechanics for the DAPs in each reconfiguration, as long as the DAPs used in a single configuration satisfy the appropriate DAP properties. Hence, makes our algorithm adaptive.

3 TREAS: A new two-round erasure-code based algorithm

In this section, we present the first, two-round, erasure-code based algorithm for implementing atomic memory service, we call TREAS. The algorithm uses $[n, k]$ MDS codes for storage. We implement and instance of the algorithm in a configuration of $n$ server processes.

The read and write operations of algorithm TREAS are implemented using $A_1$ (Alg. 10), the DAP primitives are implemented in Alg. 2 and the servers’ responses in Automaton 3. In high level, both the read and write operations take two phases to complete (similar to the ABD algorithm). As in algorithm $A_1$, a write operation $\pi$, discovers the maximum tag $t^*$ from a quorum in $Q$ by executing $c$.get-tag(); creates a new tag $t_w = \text{tag}(\pi) = (t^*, z + 1, w)$ by incorporating the writer’s own ID; and it performs a $c$.put-data($\langle t_w, v \rangle$) to propagate that pair to configuration $c$. A read operation performs $c$.get-data() to retrieve a tag-value pair, $\langle \tau, v \rangle$ form configuration $c$, and then it performs a $c$.put-data($\langle \tau, v \rangle$) to propagate that pair to the servers $c$.Servers.

To facilitate the use of erasure-codes, each server $s_i$ stores one state variable, $List$, which is a set of up to $(\delta + 1)$ (tag, coded-element) pairs. Initially the set at $s_i$ contains a single element, $List = \{(t_0, \Phi_i(v_0))\}$. Given this set we can now describe the implementation of the DAP.

A client, during the execution of a $c$.get-tag() primitive, queries all the servers in $c$.Servers for the highest tags in their Lists, and awaits responses from $\left\lceil \frac{n + k}{2} \right\rceil$ servers, with $k \geq \frac{2n}{3}$. A server upon receiving the GET-TAG request, responds to the client with the highest tag, as $\tau_{\text{max}} = \max_{(t,c) \in List} t$. Once the client receives the tags from $\left\lceil \frac{n + k}{2} \right\rceil$ servers, it selects the highest tag $t$ and returns it to $c$.

During the execution of the primitive $c$.put-data($\langle t_w, v \rangle$), a client sends the pair $(t_w, \Phi_i(v))$ to each server $s_i$. Every time a $(\text{PUT-DATA}, t_w, c_i)$ message arrives at a server $s_i$ from a writer, the pair gets added to the List. As the size of the List at each $s_i$ is bounded by $(\delta + 1)$, then following an insertion in the List, $s_i$ trims the coded-elements associated with the smallest tags. In particular,
Algorithm 2 The protocols for the DAP primitives for template $A_1$ to implement TREAS.

at each process $p_i \in \mathcal{I}$

2: procedure c.get-tag()
    send (QUERY-TAG) to each $s \in c.Servers$
4: until $p_i$ receives $(t_s, e_s)$ from $\lfloor \frac{n+n}{2} \rfloor$ servers in $c.Servers$
    $t_{max} \leftarrow \max \{ (t_s : \text{received } (t_s, v_s) \text{ from } s) \}$
6: return $t_{max}$
end procedure

8: procedure c.get-data()
    send (QUERY-LIST) to each $s \in c.Servers$
10: until $p_i$ receives $List_s$ from each server $s \in S_g$ s.t. $|S_g| \ge \lfloor \frac{n+k}{2} \rfloor$ and $S_g \subset c.Servers$
    $Tags_s^{\ge k} = \text{set of tags that appears in } k \text{ lists}$
end procedure

12: $Tags^{\ge k}_{\text{dec}} = \text{set of tags that appears in } k \text{ lists with values}$
14: $t_{max}^{\text{dec}} = \max Tags^{\ge k}_{\text{dec}}$
16: $v \leftarrow \text{decode value for } t_{max}^{\text{dec}}$
18: end procedure

Algorithm 3 The response protocols at any server $s_i \in S$ in TREAS for client requests.

at each server $s_i \in S$ in configuration $c_k$

2: State Variables:
    List \subseteq T \times C_s, \text{ initially } \{(t_0, \Phi_s(t_0))\}$

4: Upon receive (QUERY-TAG) $s_i, c_k$ from $q$
    $\tau_{max} = \max \{ t_0 : \text{List} \}$
    Send $\tau_{max}$ to $q$
6: end receive

8: Upon receive (QUERY-LIST) $s_i, c_k$ from $q$
    Send $List_s$ to $q$
10: end receive

12: If $|List| > \delta + 1$ then
    $\tau_{min} \leftarrow \min \{ t : (t, s) \in \text{List} \}$
    // remove the coded value and retain the tag
    $List \leftarrow \text{List} \setminus \{(\tau, e) : \tau = \tau_{min} \land \langle \tau, e \rangle \in \text{List} \} \cup \{(\tau_{min}, \perp)\}$
14: \hspace{1cm} return
16: Send ACK to $q$
end receive

$s_i$ replaces the coded-elements of the older tags with $\perp$, and maintains only the coded-elements associated with the $(\delta + 1)$ highest tags in the List (see Line Alg. 3[15]). The client completes the primitive operation after getting acks from $\lfloor \frac{n+k}{2} \rfloor$ servers.

A client during the execution of a c.get-data() primitive, it queries all the servers in c.Servers for their local variable List, and awaits responses from $\lfloor \frac{n+k}{2} \rfloor$ servers. Once the client receives Lists from $\lfloor \frac{n+k}{2} \rfloor$ servers, it selects the highest tag $t$, such that, (i) its corresponding value $v$ is decodable from the coded elements in the lists; and (ii) $t$ is the highest tag seen from the responses of at least $k$ Lists (see Lines Alg. 2[11][14]) and returns the pair $(t, v)$. Note that in the case where anyone of the above conditions is not satisfied the corresponding read operation does not complete.

Storage and Communication Costs for TREAS. We now briefly present the storage and communication costs associated with TREAS. Due to space limitations the proofs appear in Appendix B.

Recall that by our assumption, the storage cost counts the size (in bits) of the coded elements stored in the List variable at each server. We ignore the storage cost due to meta-data and temporary variables. Also, for the communication cost we measure the bits sent on the wire between the nodes.

Theorem 3. The TREAS algorithm has: (i) storage cost $(\delta + 1) \frac{n}{k}$, (ii) communication cost for each write at most $\frac{n}{k}$, and (iii) communication cost for each read at most to $(\delta + 2) \frac{n}{k}$.
3.1 Safety, Liveness and Performance cost of TREAS

In this section we are concerned with only one configuration \( c \), consisting of a set of servers \( S \), and a set of reader and writer clients \( R \) and \( W \), respectively. In other words, in such static system the sets \( S \), \( R \) and \( W \) are fixed, and at most \( f \leq \frac{n-k}{2} \) servers may crash fail. Below we prove Lemma 5, which proves the consistency properties of the DAP implementation of TREAS, which implies the atomicity city properties.

3.1.1 Correctness and Liveness

Now we can show that if DAP properties are satisfied from the DAP, then algorithm \( A_1 \) implements an atomic read/write algorithm. We can show that \( A_1 \) satisfy atomic guarantees and liveness if the DAP in the above algorithms satisfy the DAP consistency properties.

**Theorem 4** (Atomicity of \( A_1 \)). Suppose the DAP implementation satisfies the consistency properties \( C1 \) and \( C2 \) of Definition 31. Then any execution \( \xi \) the atomicity protocols \( A_1 \) on a configuration \( c \in C \), is atomic and live if DAPs are live in \( \xi \).

**Lemma 5.** The data-access primitives, i.e., get-tag, get-data and put-data primitives, implemented in the TREAS algorithm satisfy the consistency properties.

**Theorem 6** (Atomicity). Any execution of TREAS, is atomic.

The parameter \( \delta \) captures all the write operations that overlap with the read, until the time the reader has all data needed to attempt decoding a value. However, we ignore those write operations that might have started in the past, and never completed yet, if their tags are less than that of any write that completed before the read started. This allows us to ignore write operations due to failed writers, while counting concurrency, as long as the failed writes are followed by a successful write that completed before the read started.

**Definition 7** (Valid read operations). A read operation \( \pi \) will be called as a valid read if the associated reader remains alive until the reception of the \( \lceil \frac{n+k}{2} \rceil \) responses during the get-data phase.

**Definition 8** (Writes concurrent with a valid read). Consider a valid read operation \( \pi \). Let \( T_1 \) denote the point of initiation of \( \pi \). For \( \pi \), let \( T_2 \) denote the earliest point of time during the execution when the associated reader receives all the \( \lceil \frac{n+k}{2} \rceil \) responses. Consider the set \( \Sigma = \{ \sigma : \sigma \) is any write operation that completes before \( \pi \) is initiated\}, and let \( \sigma^* = \arg \max_{\sigma \in \Sigma} \text{tag}(\sigma) \). Next, consider the set \( \Lambda = \{ \lambda : \lambda \) is any write operation that starts before \( T_2 \) such that \( \text{tag}(\lambda) > \text{tag}(\sigma^*) \} \). We define the number of writes concurrent with the valid read operation \( \pi \) to be the cardinality of the set \( \Lambda \).

The above definition captures all the write operations that overlap with the read, until the time the reader has all data needed to attempt decoding a value. However, we ignore those write operations that might have started in the past, and never completed yet, if their tags are less than that of any write that completed before the read started. This allows us to ignore write operations due to failed writers, while counting concurrency, as long as the failed writes are followed by a successful write that completed before the read started.
Theorem 9 (Liveness). Let $\beta$ denote a well-formed execution of TREAS, with $[n, k]$, where $n$ is the number of servers and $k > n/3$, and $\delta$ be the maximum number of write operations concurrent with any valid read operation then the read and write operations $\beta$ are live.

4 Algorithm Framework for ARES

In this section, we provide the description of an atomic reconfigurable read/write storage, we call ARES. In the presentation of ARES algorithm we decouple the reconfiguration service from the shared memory emulation, by utilizing the data access primitives presented in Section 2.1. This allows ARES, to handle both the reorganization of the servers that host the data, as well as utilize a different atomic memory implementation per configuration. It is also important to note that ARES adopts a client-server architecture and separates the reader, writer and reconfiguration processes from the server processes that host the object data.

In the rest of the section we first provide the specification of the reconfiguration mechanism used in ARES, along with the properties that this service offers. Then, we discuss the implementation of read and write operations and how they utilize the reconfiguration service to ensure atomicity even in cases where read/write operations are concurrent with reconfiguration operations. The read and write operations are described in terms of the data access primitives presented in Section 2.1, and we show that if the DAP properties are satisfied then ARES preserves atomicity. This allows ARES to deploy the transformation of any atomic read/write algorithm in terms of the presented DAPs without compromising consistency.

4.1 Implementation of the Reconfiguration Service

In this section, we describe the reconfiguration service that is used in ARES, where reconfig clients introduce new configurations. In our setting, we assume throughout an execution of ARES, every configuration is attempted to be reconfigured at most once. Multiple clients may attempt concurrently to introduce a different configuration for the same index $i$ in the configuration sequence. ARES uses consensus to resolve such conflicts. In particular, each configuration $c$ is associated with an external consensus service, denoted by $c.Con$, that runs on a subset of servers in the configuration. We use the data-type $status \in \{F, P\}$, corresponding to a configuration, say $c$, to denote whether $c$ is finalized ($F$) or is still pending ($P$). Each reconfigurer may change the system configuration by introducing a new configuration identifier. The data type configuration sequence is an array of pairs $\langle cfg, status \rangle$, where $c \in C$ and $status \in \{F, P\}$. We denote each such pair by the caret over a variable name, e.g., $\hat{x}$ or $\hat{config}$, or $\hat{c}$, etc.

The service relies on an underlying sequence of configurations in the form of a “distributed list”, global configuration sequence $G_L$. In any configuration $c$, every server in $c.Servers$ has a configuration sequence variable $cseq$, initially $\langle c_0, F \rangle$, where new configurations can be added to the end of the list. We use the notation $|\hat{c}|$ to denote the length of the array. Every server in $c.Servers$ has a variable $nextC$, $nextC \in C \cup \{\perp\}$. Initially, at any server $nextC = \perp$, and once it is set to a value in $C$ it is never altered. For any $c \in C$, at any point in time, all the values of $nextC$, such that $nextC \neq \perp$, in the processes in $c.Servers$ are the same. At any point in an execution of ARES, for any $c_i, c_j \in \mathcal{C}$, we say $c_i$ points $c_j$ is a link in $G_L$ if at that point in the execution where a server in $c_j.Servers$ has $nextC = c_j$. 
The reconfiguration service consists of two major components: (i) sequence traversal, responsible for discovering the latest state of the configuration sequence $G_L$, and (ii) reconfiguration operation that installs a new configuration $G_L$.

**Algorithm 4** Sequence traversing at each process $p \in I$ of algorithm ARES.

```plaintext
procedure read-config(seq)
2: $\mu = \max\{j : seq[j].status = F\}$
3: $\hat{c} \leftarrow seq[\mu]$
4: while $\hat{c} \neq \perp$
5: \hspace{1em} $\hat{c}' \leftarrow$ read-next-config($\hat{c}, cfg$)
6: \hspace{1em} if $\hat{c}' \neq \perp$
7: \hspace{2em} $\mu \leftarrow \mu + 1$
8: \hspace{2em} put-config($seq[\mu - 1].cfg, seq[\mu]$)
9: \hspace{1em} $\hat{c} \leftarrow seq[\mu]$
10: end while
11: return $seq$
12: end procedure

procedure read-next-config($c$
14: send (READ-CONFIG) to each $s \in c.Servers$
15: until $\exists Q, P \in c.Quorums$ s.t. rec$_i$ receives nextC$_s$ from $\forall s \in Q$
16: if $\exists s \in Q$ s.t. nextC$_s$.status = $F$
17: return nextC$_s$
18: else if $\exists s \in Q$ s.t. nextC$_s$.status = $P$
19: return nextC$_s$
20: else
21: return $\perp$
22: end procedure

procedure put-config($c, nextC$)
24: send (WRITE-CONFIG, cfgPtr) to each $s \in c.Servers$
25: until $\exists Q, P \in c.Quorums$ s.t. rec$_i$ receives ACK from $\forall s \in Q$
26: end procedure
```

**Sequence Traversal.** Any read/write/reconfig operation $\pi$ utilizes the sequence traversal mechanism to discover the latest state of the global configuration sequence, as well as to ensure that such a state is discoverable by any subsequent operation $\pi'$. The sequence parsing consists of three actions: (i) get-next-config($c$), (ii) put-config($c$), and (iii) read-config($c$). We do present their specification and implementations as follows (Fig. 4):

- get-next-config($c$): The action get-next-config returns the configuration that follows $c$ in $G_L$. During get-next-config($c$) action sends READ-CONFIG messages to all the servers in $c.Servers$. Once a server receives such a message responds with the value of its nextC variable. Once it receives replies from a quorum in $c.Quorums$, then if there exists a reply that contains a nextC $\neq \perp$ the action returns nextC; otherwise it returns $\perp$.

- put-config($c, c'$): The put-config($c, c'$) action propagates $c'$ to the servers in $c.Servers$. During the action, the client sends (WRITE-CONFIG, $c'$) messages, to the servers in $c.Servers$ and waits for each server $s$ in some quorum $Q \in c.Quorums$ to respond.

- read-config($seq$): A read-config($seq$) sequentially traverses the configurations in $G_L$ in order to discover the latest state of the sequence $G_L$. At invocation, the client starts with the last finalized configuration $c_\mu$ in $seq$ (Line A412), say $c$ and enters a loop to traverse $G_L$ by invoking get-next-config($c$), which returns the next configuration, say $c'$. If $c' \neq \perp$, then: (a) $c'$ is appended at the end of the sequence $seq$; (b) a put-config($c, c_r$) is invoked to inform a quorum of servers in $c.Servers$ to update the value of their nextC variable to $c_r$; and (c) variable $c$ is set to $c_r$. If $c' = \perp$ the loop terminates the action read-config returns $seq$.

**Server Protocol.** Each server responds to requests from clients (Alg. 6). A server waits for two types of messages: READ-CONFIG and WRITE-CONFIG. When a READ-CONFIG message is received for a particular configuration $c_k$, then the server returns nextC variables of the servers in $c_k.Servers$. A WRITE-CONFIG message attempts to update the nextC$_c$ variable of the server with a particular tuple nextT$_m$. A server changes the value of its local nextC$_c$ in two cases: (i) nextC$_c.ck = \perp$, or (ii) nextC$_c.status = P$.  

11
Algorithm 5 Reconfiguration protocol of algorithm ARES.

1. at each reconfigurer \text{rec}_i
2. State Variables:
   \text{cseq}[j], \text{cseq}[i] \in C \times \{F, P\}
   with members:
4. Initializaition:
   \text{cseq}[0] = (C_0, F)
6. operation reconfig(c)
   if \(c \neq \bot\) then
   8. \text{cseq} \leftarrow \text{read-config}(\text{cseq})
   9. \text{cseq} \leftarrow \text{add-config}(\text{cseq}, c)
10. updata-config(\text{cseq})
   12. \text{cseq} \leftarrow \text{finalize-config}(\text{cseq})
14. procedure add-config(\text{seq}, c)
   16. \(d \leftarrow C_.\text{Con}.propose(c)\)
   18. \text{put-config}(c', \langle d, P \rangle)
20. return \text{seq}'
22. procedure update-config(\text{seq})
   24. \(M \leftarrow \emptyset\)
   26. for \(i = \mu : \nu\) do
   28. \(\langle \tau, v \rangle \leftarrow \text{seq}[i].cfg\).get-data()
   30. end procedure
32. procedure finalize-config(\text{seq})
   34. \text{put-config}(\text{seq}[\nu - 1].cfg, \text{seq}[\nu])
   36. return \text{seq}

Algorithm 6 Server protocol of algorithm ARES.

2. at each server \(s_i\) in configuration \(c_k\)
6. Upon receive (READ-CONFIG) \(s_i, c_k\) from \(q\)
   send nextC to \(q\)
8. end receive
10. if nextC.cfg = \bot \lor nextC.status = F then
12. send ACK to \(q\)
   end receive

Reconfiguration operation. A reconfiguration operation \text{reconfig}(c), c \in C, invoked by a non-faulty reconfiguration client \text{rec}_i, attempts to append \(c\) to \(G_L\). The operation consists of the following phases, executed consecutively by \text{rec}_i: (i) \text{read-config}; (ii) \text{add-config}; (iii) \text{update-config} and (iv) \text{finalize-config}.

\text{read-config}(\text{seq})\text{:} The phase \text{read-config}(\text{seq}) at \text{rec}_i, reads the recent global configuration sequence starting with some initial guess of \text{seq}. As described above, the \text{read-config} action completes the traversal by returning a possibly extended configuration sequence to \text{cseq}.

\text{add-config}(\text{seq}, c)\text{:} The \text{add-config}(\text{seq}, c) attempts to append a new configuration \(c\) to the end of \text{seq} (the approximation of \(G_L\)). Suppose the last configuration in \text{seq} is \(c'\), then in order to decide the most recent configuration, \text{rec}_i invokes \(C_.\text{Con}.\text{propose}(c)\), on the consensus object associated with configuration \(c'\), where \(d\) is the decided configuration identifier returned the configuration service. If \(d \neq c\), this implies that another (possibly concurrent) reconfiguration operation, invoked by a reconfigurer \text{rec}_j \neq \text{rec}_i, proposed and succeeded \(d\) as the configuration to follow \(c'\). In this case, \text{rec}_i adopts \(d\) as its own propose configuration, by adding \(\langle d, P \rangle\) to the end of its local \text{cseq} (entirely ignoring \(c\)), using the operation \text{put-config}(\langle c', \langle d, P \rangle \rangle), and returns the extended configuration \text{seq}'.

\text{update-config}(\text{seq})\text{:} Let us denote by \(\mu\) the index of the last configuration in \text{cseq}, at \text{rec}_i, such that its corresponding status is \(F\); and \(\nu\) denote the last index of \text{cseq}. Next \text{rec}_i invokes \text{update-config}(\text{seq}), which gathers the tag-value pair corresponding to the maximum tag in each of the configurations in \text{cseq}[i] for \(\mu \leq i \leq \nu\), and transfers that pair to the configuration that
Figure 1: An illustration of an execution of the reconfiguration process.

was added by the add-config action. The get-data and put-data actions are implemented with respect to the atomic algorithm that is used in each of the configurations that are accessed. Suppose \( \langle t_{max}, v_{max} \rangle \) is the tag value pair corresponding to the highest tag among the responses from all the \( \nu - \mu + 1 \) configurations. Then, \( \langle t_{max}, v_{max} \rangle \) is written to the configuration \( d \) via the invocation of \( \text{cfg}.\text{put-data}(\langle t_{max}, v_{max} \rangle) \).

**finalize-config(cseq)**: Once the tag-value pair is transferred, in the last phase of the reconfiguration operation, \( rec_i \) executes finalize-config(cseq), to update the status of the last configuration in cseq, i.e. \( d=cseq[\nu] \), to \( F \). \( rec_i \) informs a quorum of servers in the previous configuration, i.e. in some \( Q \in c.\text{Quorums} \), about the change of status, by executing the \( c.\text{put-config}(c, \langle d,F \rangle) \) action.

Fig. 1 illustrates an example execution of a reconfiguration operation recon \( (c_5) \). In this example, the reconfigurer \( rec_i \) goes through a number of configuration queries (read-next-config) before it reaches configuration \( c_4 \) in which a quorum of servers replies with \( \text{nextC.cfg} = \bot \). There it proposes \( c_5 \) to the consensus object of \( c_4 \) (\( c_4.\text{Con.propose}(c_5) \) on arrow 10), and once \( c_5 \) is decided, recon \( (c_5) \) completes after executing finalize-config \( (c_5) \).

**Read and Write operations.** The read and write operations in ARES are expressed in terms of the DAP primitives (see Section 3). A read consists of an execution of get-data primitive followed by a put-data primitive, while a write consists of calls to get-tag and put-data primitives. This provides the flexibility to ARES to use different implementation of DAP primitives, without changing the ARES framework. At a high-level, a write (or read) operation is executed where the client: (i) obtains the latest configuration sequence by using the read-config action of the reconfiguration service, (ii) queries the configurations, in cseq, starting from the last finalized configuration to the end of the discovered configuration sequence, for the latest \( \langle \text{tag}, \text{value} \rangle \) pair with a help of get-tag (or get-data) operation, and (iii) repeatedly propagates a new \( \langle \text{tag}, \text{value} \rangle \) pair (the largest \( \langle \text{tag}, \text{value} \rangle \) pair) with put-data in the last configuration of its local sequence, until no additional configuration is observed. Now we describe the execution of a read or write operation \( \pi \) in more detail.

In line Alg. 7:8 for the writer (or line Alg. 7:31 for reader), a write (or read) operation is invoked at a client \( p \), then \( p \) issues a read-config action to obtain the latest introduced configuration in \( G_L \).

In lines Alg. 7:32 if \( \pi \) is a write (resp. Alg. 7:33 if \( \pi \) is a read), \( p \) detects the last finalized entry in
Algorithm 7 Write and Read protocols at the clients for ARES.

<table>
<thead>
<tr>
<th>Line</th>
<th>Code</th>
</tr>
</thead>
<tbody>
<tr>
<td>2</td>
<td>Write Operation:</td>
</tr>
<tr>
<td></td>
<td>at each writer $w_i$</td>
</tr>
<tr>
<td>4</td>
<td>State Variables:</td>
</tr>
<tr>
<td></td>
<td>$cseq[j] \in C \times {F, P}$ with members:</td>
</tr>
<tr>
<td>6</td>
<td>Initialization:</td>
</tr>
<tr>
<td></td>
<td>$cseq[0] = ({0}, F)$</td>
</tr>
<tr>
<td>8</td>
<td>operation write(val), val $\in V$</td>
</tr>
<tr>
<td></td>
<td>$cseq \leftarrow \text{read-config}(cseq)$</td>
</tr>
<tr>
<td></td>
<td>$\mu \leftarrow \max((i : cseq[i].\text{status} = F))$</td>
</tr>
<tr>
<td>10</td>
<td>$\nu \leftarrow</td>
</tr>
<tr>
<td></td>
<td>for $i = \mu : \nu$ do</td>
</tr>
<tr>
<td></td>
<td>$\tau_{\max} \leftarrow \max(cseq[i].cfg.\text{get-tag}(), \tau_{\max})$</td>
</tr>
<tr>
<td></td>
<td>$(\tau, \nu) \leftarrow (\tau_{\max}, \nu, |val|)$</td>
</tr>
<tr>
<td></td>
<td>done $\leftarrow$ false</td>
</tr>
<tr>
<td></td>
<td>while not done do</td>
</tr>
<tr>
<td></td>
<td>$cseq[\nu].cfg.\text{put-data}((\tau, \nu))$</td>
</tr>
<tr>
<td></td>
<td>$cseq \leftarrow \text{read-config}(cseq)$</td>
</tr>
<tr>
<td>18</td>
<td>if $</td>
</tr>
<tr>
<td></td>
<td>done $\leftarrow$ true</td>
</tr>
<tr>
<td></td>
<td>else</td>
</tr>
<tr>
<td></td>
<td>$\nu \leftarrow</td>
</tr>
<tr>
<td></td>
<td>end while</td>
</tr>
<tr>
<td>22</td>
<td>end operation</td>
</tr>
<tr>
<td>24</td>
<td>Read Operation:</td>
</tr>
<tr>
<td></td>
<td>at each reader $r_j$</td>
</tr>
<tr>
<td>26</td>
<td>State Variables:</td>
</tr>
<tr>
<td></td>
<td>$cseq[j] \in C \times {F, P}$ with members:</td>
</tr>
<tr>
<td>28</td>
<td>Initialization:</td>
</tr>
<tr>
<td></td>
<td>$cseq[0] = ({0}, F)$</td>
</tr>
<tr>
<td>30</td>
<td>operation read( )</td>
</tr>
<tr>
<td></td>
<td>$cseq \leftarrow \text{read-config}(cseq)$</td>
</tr>
<tr>
<td></td>
<td>$\mu \leftarrow \max((j : cseq[j].\text{status} = F))$</td>
</tr>
<tr>
<td></td>
<td>$\nu \leftarrow</td>
</tr>
<tr>
<td></td>
<td>for $i = \mu : \nu$ do</td>
</tr>
<tr>
<td></td>
<td>$(\tau, \nu) \leftarrow \max(cseq[i].cfg.\text{get-data}(), (\tau, \nu))$</td>
</tr>
<tr>
<td></td>
<td>done $\leftarrow$ false</td>
</tr>
<tr>
<td></td>
<td>while not done do</td>
</tr>
<tr>
<td></td>
<td>$cseq[\nu].cfg.\text{put-data}((\tau, \nu))$</td>
</tr>
<tr>
<td></td>
<td>$cseq \leftarrow \text{read-config}(cseq)$</td>
</tr>
<tr>
<td>40</td>
<td>if $</td>
</tr>
<tr>
<td></td>
<td>done $\leftarrow$ true</td>
</tr>
<tr>
<td>42</td>
<td>else</td>
</tr>
<tr>
<td></td>
<td>$\nu \leftarrow</td>
</tr>
<tr>
<td>44</td>
<td>end while</td>
</tr>
<tr>
<td></td>
<td>end operation</td>
</tr>
</tbody>
</table>

$cseq$, say $\mu$, and performs a $cseq[j].\text{conf.\text{get-tag}()}$ action if $\pi$ is a write, or $cseq[j].\text{conf.\text{get-data}()}$ action if $\pi$ is a read action, for $\mu \leq j \leq |cseq|$. Then $p$ discovers the maximum tag among all the returned tags ($\tau_{\max}$) or tag-value pairs ($\langle \tau_{\max}, v_{\max} \rangle$) respectively. If $\pi$ is a write, $p$ increments the maximum tag discovered (by incrementing the integer part of $\tau_{\max}$), generates a new tag, say $\tau_{\new}$, and assigns $(\tau, \nu)$ to $\langle \tau_{\new}, \|val\| \rangle$, where $val$ is the value he wants to write (Line Alg. 7[13]). If $\pi$ is a read, then $p$ assigns $(\tau, \nu)$ to $\langle \tau_{\max}, v_{\max} \rangle$, i.e., the maximum discovered tag-value pair.

In lines Alg. 7[15][21] if $\pi$ is a write, or lines Alg. 7[37][43] if $\pi$ is a read, $p$ repeatedly executes the $cseq[\nu].cfg.\text{put-data}((\tau, \nu))$ action, where $\nu = |cseq|$, followed by executing read-config action, to examine whether new configurations were introduced in $G_L$. Let $cseq'$ be the sequence returned by the read-config action. If $|cseq'| = |cseq|$ then no new configuration is introduced, and the read/write operation terminates; otherwise, $p$ sets $cseq$ to $cseq'$ and repeats the two actions. Note, in an execution of ARES, two consecutive read-config operations that return $cseq'$ and $cseq''$ respectively must hold that $cseq'$ is a prefix of $cseq''$, and hence $|cseq'| = |cseq''|$ only if $cseq' = cseq''$.

4.2 Properties of Reconfiguration

Notations and definitions. For a server $s$, we use the notation $s.var|_{\sigma}$ to refer to the value of the state variable $var$, in $s$, at a state $\sigma$ of an execution $\xi$. If server $s$ crashes at a state $\sigma_j$ in an execution $\xi$ then $s.var|_{\sigma} \triangleq s.var|_{\sigma_j}$ for any state variable $var$ and for any state $\sigma$ that appears after $\sigma_j$ in $\xi$.

**Definition 10** (Tag of a configuration). Let $c \in C$ be a configuration, $\sigma$ be a state in some execution $\xi$ then we define the tag of $c$ at state $\sigma$ as $\text{tag}(c)|_{\sigma} \triangleq \min_{\xi \in \text{Qorums}} \max_{s \in \text{Qorums}} (s.tag|_{\sigma})$. We drop the suffix $|_{\sigma}$, and simply denote as $\text{tag}(c)$, when the state is clear from the context.
Definition 11. Let \( c^p_i = p.cseq|_{\sigma} \). Then we define as \( \mu(c^p_i) \equiv \max\{i : c^p_i[i].status = F\} \) and \( \nu(c^p_i) \equiv |c^p_i| \), where \( |c^p_i| \) is the number of elements in configuration vector \( c^p_i \) that are not equal to \( \perp \).

Definition 12 (Prefix order). Let \( x \) and \( y \) be any two configuration sequences. We say that \( x \) is a prefix of \( y \), denoted by \( x \preceq_p y \), if \( x[j].cfg = y[j].cfg \), for all \( j \) such that \( x[j] \neq \perp \).

Next we analyze some properties of ARES. The first lemma shows that any two configuration sequences have the same configuration identifiers in the same indexes.

Lemma 13 (Configuration Uniqueness). For any processes \( p, q \in I \) and any states \( \sigma_1, \sigma_2 \) in an execution \( \xi \), it must hold that \( c^p_{\sigma_1}[i].cfg = c^q_{\sigma_2}[i].cfg \), \( \forall i \) s.t. \( c^p_{\sigma_1}[i].cfg, c^q_{\sigma_2}[i].cfg \neq \perp \).

We can now move to an important lemma that shows that any read-config action returns an extension of the configuration sequence returned by any previous read-config action. First, we show that the last finalized configuration observed by any read-config action is at least as recent as the finalized configuration observed by any subsequent read-config action.

Lemma 14 (Configuration Prefix). Let \( \pi_1 \) and \( \pi_2 \) two completed read-config actions invoked by processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \) in an execution \( \xi \). Let \( \sigma_1 \) be the state after the response step of \( \pi_1 \) and \( \sigma_2 \) the state after the response step of \( \pi_2 \). Then \( c^p_{\sigma_1} \preceq_p c^p_{\sigma_2} \).

Thus far we focused on the configuration member of each element in \( cseq \). As operations do get in account the status of a configuration, i.e. \( P \) or \( F \), in the next lemma we will examine the relationship of the last finalized configuration as detected by two operations. First we present a lemma that shows the monotonicity of the finalized configurations.

Lemma 15 (Configuration Progress). Let \( \pi_1 \) and \( \pi_2 \) two completed read-config actions invoked by processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \) in an execution \( \xi \). Let \( \sigma_1 \) be the state after the response step of \( \pi_1 \) and \( \sigma_2 \) the state after the response step of \( \pi_2 \). Then \( \mu(c^p_{\sigma_1}) \leq \mu(c^p_{\sigma_2}) \).

Theorem 16. Let \( \pi_1 \) and \( \pi_2 \) two completed read-config actions invoked by processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \) in an execution \( \xi \). Let \( \sigma_1 \) be the state after the response step of \( \pi_1 \) and \( \sigma_2 \) the state after the response step of \( \pi_2 \). Then the following properties hold: (a) \( c^p_{\sigma_2}[i].cfg = c^p_{\sigma_1}[i].cfg \), for \( 1 \leq i \leq \nu(c^p_{\sigma_1}) \), (b) \( c^p_{\sigma_1} \preceq_p c^p_{\sigma_2} \), and (c) \( \mu(c^p_{\sigma_1}) \leq \mu(c^p_{\sigma_2}) \)

4.3 ARES Safety

Once we showed some properties that are satisfied by the reconfiguration service in any execution, we can now proceed to examine whether our algorithm satisfies the safety (atomicity) conditions. The propagation of the information of the distributed object is achieved using the get-tag, get-data, and put-data actions. We assume that the primitives used satisfy properties C1 and C2 as presented in Section 2.1, and we will show that, given such assumption, ARES satisfies atomicity.

We begin with a lemma that states that if a reconfiguration operation retrieves a configuration sequence of length \( k \) during its read-config action, then it installs/finalizes the \( k + 1 \) configuration in the global configuration sequence \( G_L \).
Lemma 17. Let \( \pi \) be a complete reconfiguration operation by a reconfigurer \( rc \) in an execution \( \xi \) of ARES. If \( \sigma_1 \) is the state in \( \xi \) following the termination of the read-config action during \( \pi \), then \( \pi \) invokes a finalize-config(\( c^{rc}_{\sigma_2} \)) at a state \( \sigma_2 \) in \( \xi \), with \( \nu(c^{rc}_{\sigma_2}) = \nu(c^{rc}_{\sigma_1}) + 1 \).

The next lemma states that each finalized configuration \( c \) at index \( j \) in a configuration sequence \( p.cseq \) at any process \( p \), is finalized by some reconfiguration operation \( \rho \). To finalize \( c \), the lemma shows that \( \rho \) must obtain a configuration sequence such that its last finalized configuration that appears before \( c \) in the configuration sequence \( p.cseq \). In other words, reconfigurations always finalize configurations that are ahead from their latest observed final configuration, and it seems like “jumping” from one final configuration to the next.

Lemma 18. Suppose \( \xi \) is an execution of ARES. For any state \( \sigma \) in \( \xi \), if \( c^p_j.\text{status} = F \) for some process \( p \in I \), then there exists a reconfig action \( \rho \) by a reconfigurer \( rc \in G \), such that (i) \( rc \) invokes finalize-config(\( c^{rc}_{\sigma} \)) during \( \rho \) at some state \( \sigma' \) in \( \xi \), (ii) \( \nu(c^{rc}_{\sigma}) = j \), and (iii) \( \mu(c^{rc}_{\sigma}) < j \).

In ARES, before a read/write/reconfig completes it propagates the maximum tag it discovered by executing the put-data action in the last configuration of its local configuration sequence. When a subsequent operation is invoked, reads the latest configuration sequence and, beginning from the last finalized configuration, it invokes read-data to all the configurations until the end of the sequence. The lemma shows that the latter operation will retrieve a tag which is higher than the tag used in the put-data action. For the following proof we use the notation \( \nu'(c^p_j) = c^p_j[\nu(c^p_j)].cfg \). In other words, \( \nu'(c^p_j) \) is the last configuration in the sequence \( c^p_j \).

Lemma 19. Let \( \pi_1 \) and \( \pi_2 \) be two completed read/write/reconfig operations invoked by processes \( p_1 \) and \( p_2 \) in \( I \), in an execution \( \xi \) of ARES, such that, \( \pi_1 \rightarrow \pi_2 \). If \( c_{p_1}.\text{put-data}(\langle \tau_{\pi_1}, v_{\pi_1} \rangle) \) is the last put-data action of \( \pi_1 \) and \( \sigma_2 \) is the state in \( \xi \), after the completion of the first read-config action of \( \pi_2 \), then there exists a \( c_{p_2}.\text{put-data}(\langle \tau, \nu \rangle) \) action in some configuration \( c_2 = c_{\sigma_2}[\nu].cfg \), for \( \mu(c^{p_2}_{\sigma_2}) \leq i \leq \nu(c^{p_2}_{\sigma_2}) \), such that (i) it completes in a state \( \sigma' \) before \( \sigma_2 \) in \( \xi \), and (ii) \( \tau \geq \tau_{\pi_1} \).

The following lemma shows the consistency of operations as long as the DAP used satisfy properties C1 and C2.

Lemma 20. Let \( \pi_1 \) and \( \pi_2 \) denote completed read/write operations in an execution \( \xi \), from processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \). If \( \tau_{\pi_1} \) and \( \tau_{\pi_2} \) are the local tags at \( p_1 \) and \( p_2 \) after the completion of \( \pi_1 \) and \( \pi_2 \) respectively, then \( \tau_{\pi_1} \leq \tau_{\pi_2} \); if \( \pi_1 \) is a write then \( \tau_{\pi_1} < \tau_{\pi_2} \).

And the safety result of this section follows as:

Theorem 21 (Atomicity). ARES implements a reconfigurable atomic storage service, given that the get-data, get-tag, and put-data primitives used satisfy properties C1 and C2 of Definition [37].

In ARES, each configuration may utilize a separate way of implementing the DAP primitives as stated below:

Remark 22. Algorithm ARES satisfies atomicity even when the DAP primitives used in two different configurations \( c_1 \) and \( c_2 \) are not the same, given that the \( c_i \), get-tag, \( c_i \), get-data, and the \( c_i \), put-data primitives used in each \( c_i \) satisfy properties C1 and C2 of Definition [37].
4.4 Latency Analysis

No liveness properties are provided for ARES, this is because our reconfiguration mechanism uses consensus, therefore, we provide a conditional performance analysis here. In this section, we examine closely the latencies of each operation in ARES, and study the completion of each operation under various environmental conditions. For our analysis, we assume that local computation take negligible time and delays are only introduced due to the message exchange among the processes. We measure delays in time units of some global clock $T$, which is visible only to an external viewer. No process has access to $T$ and the clock. Let $d$ and $D$ be the minimum and maximum durations taken by any message to go from one process to another. Also, let $T(\pi)$ denote the communication delay of an operation (or action) $\pi$. For simplicity of our analysis, we assume that any propose operation to a consensus instance terminates in $T(CN)$ time units. Given $d$ and $D$, and through inspection of the algorithm we can provide the delay bounds of the operations and actions used by ARES as follows:

**Lemma 23.** If any message send from a process $p_1$ to a process $p_2$, s.t. $p_1, p_2 \in I \cup S$, takes at least $d$ and at most $D$ time units to be delivered, then the following operations may terminate within the following time intervals: (i) $2d \leq T(\text{put-config}) \leq 2D$ and (ii) $2d \leq T(\text{read-next-config}) \leq 2D$.

From Lemma 23 we can derive the delay of a read-config action.

**Lemma 24.** For any read-config action $\phi$ such that it accepts an input seq and returns seq', if $\mu = \mu(\text{seq})$ and $\nu = \nu(\text{seq'})$ then for $\phi$ $4d(\nu - \mu + 1) \leq T(\phi) \leq 4D(\nu - \mu + 1)$.

**Lemma 25.** Starting from the last state of $\xi$, $\sigma$, and given that $d$ is the minimum communication delay, then $k$ configurations can be appended to $c_{\sigma}$, in time: $T(k) \geq 4d \sum_{i=1}^{k} i + k(T(CN) + 2d)$ in our execution construction.

Now, we can bound the latency of that a read or write operation. The implementation, and the maximum delays, of the DAPs used by ARES impact the delay of read and write operations. For our analysis, taking ABD algorithm as an example, we assume that the implementation of each of get-data, get-tag and put-data has two phases of communication.

**Lemma 26.** If any message send from a process $p_1$ to a process $p_2$, s.t. $p_1, p_2 \in I \cup S$, takes at least $d$ and at most $D$ time units to be delivered, then the DAPs may terminate in: (i) $2d \leq T(\text{put-data}) \leq 2D$; (ii) $2d \leq T(\text{get-tag}) \leq 2D$; and (iii) $2d \leq T(\text{get-data}) \leq 2D$.

Having the delays for the DAPs we can now compute the delay of a read/write operation $\pi$.

**Lemma 27.** Let $\sigma_s$ and $\sigma_e$ be the states before the invocation and after the completion step of a read/write operation $\pi$ by $p$ respectively, in some execution $\xi$ of ARES. Then $\pi$ takes time at most: $T(\pi) \leq 6D \left[ \nu(\text{c}_{\sigma_e}) - \mu(\text{c}_{\sigma_s}) + 2 \right]$.

It remains now to examine if a read/write operation may “catch up” with any ongoing recon-figurations. For worst case analysis, we assume that reconfiguration operations may communicate respecting the minimum delay $d$, whereas read and write operations suffer the maximum delay $D$ in each message exchange. We consider three cases with respect to the number of configurations installed $k$, and the minimum delay $d$ of messages: (i) $k$ is finite, and $d$ may be very small; (ii) $k$ is infinite, and $d$ may be very small; (iii) $k$ is infinite, and $d$ can be bounded.
\( k \) is finite, and \( d \) may be unbounded small. As the \( d \) is unbounded, it follows that reconfigurations may be installed almost instantaneously. Let us first examine what is the maximum delay bound of a any read/write operation. [NN:There is something wrong with this statement.]

\( k \) is infinite, and \( d \) is bounded. We show bounds on \( d \) with respect to the \( D \) and \( k \) if we want to allow a read/write operation to reach ongoing reconfigurations.

**Lemma 28.** A read/write operation \( \pi \) may terminate in any execution \( \xi \) of ARES given that \( k \) configurations are installed during \( \pi \), if \( d \geq \frac{3D}{k} - \frac{T(CN)}{2(k+2)} \)

5 Efficient state transfer during reconfiguration

In this section, we show that TREAS can be adapted to allow reconfiguration where the object values are transferred directly from the servers in configuration, that is already finalized, to those of a new configuration, without the recon client handling object values. In Algs. 8 and 9, we show the changes necessary to adapt ARES and TREAS for achieve this. Here every configuration uses TREAS as the underlying atomic memory emulation algorithm, and we refer to this algorithm as ARES-TREAS.

In ARES, the procedure update-config (see Alg. 5) is modified as shown in Alg. 8. Consider a reconfiguration client \( rc \), which invokes update-config, during a reconfiguration operation, where it iteratively gathers the tag-config ID pairs in the set variable \( M \) by calling get-tag (lines Alg. 8: 5-9). Suppose \( \langle \tau, C \rangle \) is the tag and configuration ID pair corresponding to the highest tag in \( M \) (lines Alg. 8: 11). Next, \( rc \) executes procedure forward-code-element(REQ-FW-CODE-ELEM, \( \tau, C, C' \)), to send a requests to the servers in \( C \) to forward their respective coded elements corresponding to \( \tau \), to each server in \( C'.Servers \). Suppose the MDS code parameters in \( C \) and \( C' \) are \([n, k]\) and \([n', k']\), respectively, such that, \( |C.Servers| = n, |C'.Servers| = n', \) and for some \( k \geq \frac{2n}{3} \) and \( k' \geq \frac{2n'}{3} \). In forward-code-element, the call to md-primitive(REQ-FW-CODE-ELEM, \( \tau C' \)), presented in [21], delivers the message (REQ-FW-CODE-ELEM, \( \tau C' \)) to either every non-faulty servers in \( C.Servers \) or none. We rely on the semantics of md-primitive to avoid lingering of coded elements for ever in \( D \) due to crash failure of the \( rc \), or servers in \( C.Servers \). For example, suppose \( rc \) communicates only to one server, say \( s_i \), in \( C.Servers \) and crashes, then the rest of the servers in \( C \) would not send their coded elements to the servers in \( C' \). As a result, the coded element from \( s_i \) will linger around in the \( D \) variables in the servers in \( C'.Servers \) without ever being removed, thereby, progressively increasing the storage cost. Upon delivering these messages to any server \( s_i \), in \( C.Servers \), if \( \langle \tau, e_i \rangle \) in \( List \) in \( s_i \), then \( s_i \) sends (FWD-CODE-ELEM, \( \langle \tau, e_i \rangle, rc \)) to servers in \( C' \).

Next upon receiving any of the FWD-CODE-ELEM messages, at any server \( s'_j \) in \( C'.Servers \) if \( rc \in Recons \) in \( s'_j \) (Alg. 9: 10) then it ignores it because \( rc \) has already been updated by \( s'_j \), regarding the object value of \( \tau \). Otherwise, \( s'_j \) checks if \( \langle \tau, e_j \rangle \in List \) (Alg. 9: 11), if it is not, then \( s'_j \) adds the incoming pair \( \langle \tau, e_i \rangle \) to \( D \). Next, \( s'_j \) checks if the value for \( \tau \) is decodable (Alg. 9: 12), from the coded elements in \( D \), if it is, then \( s'_j \) decodes the value \( v \), using decoder for \( C' \), with parameters \([n, k]\) and re-encodes, according to parameters \([n', k']\), to get \( e'_j \equiv \Phi_{C'}(v) \). Then \( s'_j \) proceeds to store \( \langle \tau, e'_j \rangle \) in a similar steps as in the put-data response in the TREAS (Alg. 9). Then in lines Alg. 9: 20-22, if \( \langle \tau, * \rangle \in List \) then \( s'_j \) adds \( rc \) to the list \( Recons \) and \( s'_j \) sends \( rc \) an ACK. Finally, once \( rc \) receives ACKs from \( \left\lceil \frac{n'+k'}{2} \right\rceil \) servers in \( C'.Servers \) it completes the call to update-config.
Finally, it can be shown that ARES-TREAS implements an atomic memory service as stated in the following theorem.

**Algorithm 8** Alternate update-config for the reconfiguration protocol of ARES.

```plaintext
procedure update-config(seq)
2: \( \mu \leftarrow \max\{j : seq[j].status = F\} \)
3: \( \nu \leftarrow |seq| \)
4: \( M \leftarrow \emptyset \)
5: for \( i = 0 \) to \( \nu \) do
6: \( \tau \leftarrow seq[i].cfg.get\_data() \)
7: \( \tau \leftarrow seq[i].cfg.get\_tag() \)
8: \( M \leftarrow M \cup \{\tau\} \)
9: \( M \leftarrow M \cup \{\tau, C\} \)
10: \( \tau \leftarrow \max\{\tau, C\} \)
11: \( \nu \leftarrow |seq| \)
12: \( C \leftarrow put-data(\tau, C') \)
13: \( forward-code-element(\tau, C') \)
14: end procedure

procedure forward-code-element(\( \tau, C' \))
15: Call md-primitive (REQ-FW-CODE-ELEM, \( \tau, C' \)) on servers in \( C \)
16: until ACK from \( \lceil \frac{n + k'}{2} \rceil \) servers in \( C'.Servers \)
```

**Algorithm 9** Additional server protocol and state-variable at a server in TREAS.

```plaintext
at each server \( s_i \) in any configuration
2: Additional State Variables:
3: \( D \leftarrow D \cup \{\tau, e_i\} \)
4: \( f^{\#} s_i \) in configuration \( C \) *
5: Upon recv (REQ-FW-CODE-ELEM, \( t, C' \)) \( s_i \) from rc
6: if \( (t, e_i) \in List \) then
7: Send (FWD-CODE-ELEM, \( \tau, e_i, rc \)) to servers in \( C' \)
8: if \( (t, e_i) \in List \) then
9: \( D \leftarrow D \cup \{\tau, e_i\} \)
10: if \( (t, e_i) \in List \) then
```

**Theorem 29** (Atomicity). Algorithm ARES-TREAS implements a reconfigurable atomic storage service, if get-data, get-tag, and put-data primitives used satisfy C1 and C2 of Definition [31]

6 Conclusions

In this paper, we presented an new algorithmic framework suitable for reconfiguration of the set of servers that implements erasure code-based atomic memory service in message-passing environments. We also provided a new two-round erasure code-based algorithm that has near optimal storage cost, and bandwidth costs per read or write operation. Moreover, this algorithm is suitable specifically where during new configuration installation the object values passes directly...
from servers in older configuration to those in the newer configurations. Future work will involve adding efficient repair and reconfiguration using regenerating codes.

References


Appendix

A The data access primitives

In this section we focus on algorithms that utilize logical tags to implement atomic read/write objects, and analyze their correctness and liveness in terms of three data access primitives (DAP). These data access primitives are specific to any configuration $c$ in context: (i) put-data($\langle \tau, v \rangle$), (ii) $c$.get-data(), and (iii) $c$.get-tag(). Assuming a set of totally ordered timestamps $T$, a value domain of the distributed atomic object $V$, and a set of configuration identifiers $C$, the three primitives can be defined over a configuration $c \in C$, tag $\tau \in T$, and a value $v \in V$ as follows:

Definition 30 (Data Access Primitives). Given a configuration identifier $c \in C$, any non-faulty client process $p$ may invoke the following data access primitives during an execution $\xi$:

D1. $c$.get-tag(): returns a tag $\tau \in T$

D2. $c$.get-data(): returns a tag-value pair $(\tau, v) \in T \times V$

D3. $c$.put-data($\langle \tau, v \rangle$): the tag-value pair $(\tau, v) \in T \times V$ as argument

where $c$ is added to specify the configuration specific implementation of these primitives.

Most logical timestamp-based shared atomic memory implementations, assume that a tag $\tau \in T$ is defined as a pair $(z, w)$, where $z \in \mathbb{N}$ and $w \in W$, an ID of a writer. Notice that tags could be defined in any totally ordered domain and given that this domain is countably infinite, then there can be a direct mapping to the domain we assume. For any $\tau_1, \tau_2 \in T$ we define $\tau_2 > \tau_1$ if (i) $\tau_2.z > \tau_1.z$ or (ii) $\tau_2.z = \tau_1.z$ and $\tau_2.w > \tau_1.w$. Now consider an algorithmic template (see
Algorithm 10 Read and write operations of generic algorithm $A_1$

```
operation read()
2:  \langle t, v \rangle \leftarrow c.get-data()
    c.put-data(\langle t, v \rangle)
4:  return \langle t, v \rangle
end operation
```

6:  operation write(v)
    \( t \leftarrow c.get-tag() \)
8:  \( t_w \leftarrow (t.z + 1, w) \)
    c.put-data(\langle t_w, v \rangle)
10: end operation

Automaton [10], we call $A_1$. In brief, a read operation in $A_1$ performs $c.c.get-data()$ to retrieve a tag-value pair, $\langle \tau, v \rangle$ form configuration $c$, and then it performs a $c.c.put-data(\langle \tau, v \rangle)$ to propagate that pair to configuration $c$. A write operation is similar to the read but before performing the put-data action it generates a new tag which associates with the value to be written. We can show that $A_1$ satisfy atomic guarantees and liveness if the DAP in the above algorithms satisfy the following consistency properties:

**Definition 31 (DAP Consistency Properties).** In an execution $\xi$ we say that a DAP operation in an execution $\xi$ is complete if both the invocation and the matching response step appear in $\xi$. If $\Pi$ is the set of complete DAP operations in execution $\xi$ then for any $\phi, \pi \in \Pi$:

1. **C1** If $\phi$ is a $c.c.put-data(\langle \tau_\phi, v_\phi \rangle)$, for $c \in C$, $\tau_1 \in T$ and $v_1 \in V$, and $\pi$ is a $c.c.get-tag()$ (or a $c.c.get-data()$) that returns $\tau_\pi \in T$ (or $\langle \tau_\pi, v_\pi \rangle \in T \times V$) and $\phi$ completes before $\pi$ in $\xi$, then $\tau_\pi \geq \tau_\phi$.

2. **C2** If $\phi$ is a $c.c.get-data()$ that returns $\langle \tau_\pi, v_\pi \rangle \in T \times V$, then there exists $\pi$ such that $\pi$ is $c.c.put-data(\langle \tau_\pi, v_\pi \rangle)$ and $\phi$ did not complete before the invocation of $\pi$, and if no such $\pi$ exists in $\xi$, then $\langle \tau_\pi, v_\pi \rangle$ is equal to $\langle t_0, v_0 \rangle$.

Algorithm 11 Read and write operations of generic algorithm $A_2$

```
operation read()
2:  \langle t, v \rangle \leftarrow c.c.get-data()
    return \langle t, v \rangle
end operation
```

```
operation write(v)
6:  \( t \leftarrow c.c.get-tag() \)
8:  \( t_w \leftarrow (t.z + 1, w) \)
    c.c.put-data(\langle t_w, v \rangle)
end operation
```

A slightly different algorithmic template $A_2$ (see Automaton [11]), captures algorithms where read operations avoid the propagation phase. In $A_2$ the write protocol is the same as in $A_1$ however the read operation terminates as soon as the get-data action returns. In addition to C1 and C2, algorithms that are transformed in the generic algorithm $A2$ need to satisfy the following property:

3. **C3** If $\phi$ is a $c.c.get-data()$ that returns $\langle \tau_\phi, v_\phi \rangle$ and $\pi$ is a $c.c.get-data()$ that returns $\langle \tau_\pi, v_\pi \rangle$, and $\phi \rightarrow \pi$, then $\tau_\phi \leq \tau_\pi$.

Now we can show that if those properties are satisfied from the DAP, then algorithms $A_1$ and $A_2$ implement an atomic read/write algorithm.
**Theorem 32** (Atomicity of $A_1$). Suppose the DAP implementation satisfies the consistency properties $C1$ and $C2$ of Definition [3][7]. Then any execution $\xi$ the atomicity protocols $A_1$ on a configuration $c \in C$, is atomic and live if DAPs are live in $\xi$.

**Proof.** We prove atomicity by proving properties $A1$, $A2$ and $A3$ appearing in the definition of atomicity in Section 2 for any execution of the algorithm.

Property $A1$: Consider two operations $\phi$ and $\pi$ such that $\phi$ completes before $\pi$ is invoked. We need to show that it cannot be the case that $\pi \preceq \phi$. We break our analysis into the following four cases:

Case (a): Both $\phi$ and $\pi$ are writes. The $c$.put-data($*$) of $\phi$ completes before $\pi$ is invoked. By property $C1$ the tag $\tau_\pi$ returned by the $c$.get-data() at $\pi$ is at least as large as $\tau_\phi$. Now, since $\tau_\pi$ is incremented by the write operation then $\pi$ puts a tag $\tau'_\pi$ such that $\tau_\phi < \tau'_\pi$ and hence we cannot have $\pi \preceq \phi$.

Case (b): $\phi$ is a write and $\pi$ is a read. In execution $\xi$ since $c$.put-data($\langle t_\phi, * \rangle$) of $\phi$ completes before the $c$.get-data() of $\pi$ is invoked, by property $C1$ the tag $\tau_\pi$ obtained from the above $c$.get-data() is at least as large as $\tau_\phi$. Now $\tau_\phi \leq \tau_\pi$ implies that we cannot have $\pi \preceq \phi$.

Case (c): $\phi$ is a read and $\pi$ is a write. Let the id of the writer that invokes $\pi$ we $w_\pi$. The $c$.put-data($\langle t_\phi, * \rangle$) call of $\phi$ completes before $c$.get-tag() of $\pi$ is initiated. Therefore, by property $C1$ get-tag(c) returns $\tau$ such that, $\tau_\phi \leq \tau$. Since $\tau_\pi$ is equal to $(\tau.z + 1, w_\pi)$ by design of the algorithm, hence $\tau_\pi > \tau_\phi$ and we cannot have $\pi \preceq \phi$.

Case (d): Both $\phi$ and $\pi$ are reads. In execution $\xi$ the $c$.put-data($\langle t_\phi, * \rangle$) is executed as a part of $\phi$ and completes before $c$.get-data() is called in $\pi$. By property $C1$ of the data-primitives, we have $\tau_\phi \leq \tau_\pi$ and hence we cannot have $\pi \preceq \phi$.

Property $A2$: Note that because $T$ is well-ordered we can show that this property by first showing that every write has a unique tag. This means any two pair of writes can be ordered. Now, a read can be ordered . Note that a read can be ordered w.r.t. to any write operation trivially if the respective tags are different, and by definition, if the tags are equal the write is ordered before the read.

Now observe that two tags generated from two write operations from different writers are necessarily distinct because of the id component of the tag. Now if the operations, say $\phi$ and $\pi$ are writes from the same writer then by well-formedness property the second operation is invoked after the first completes, say without loss of generality $\phi$ completes before $\pi$ is invoked. In that case the integer part of the tag of $\pi$ is higher by property $C1$, and since the $c$.get-tag() is followed by $c$.put-data($*$). Hence $\pi$ is ordered after $\phi$.

Property $A3$: This is clear because the tag of a reader is defined by the tag of the value it returns by property (b). Therefore, the reader’s immediate previous value it returns. On the other hand if does note return any write operation’s value it must return $v_0$. 

**Theorem 33** (Atomicity of $A_2$). Suppose the DAP implementation satisfies the consistency properties $C1$, $C2$ and $C3$ of Definition [3][7]. Then any execution $\xi$ the atomicity protocols $A_2$ on a configuration $c \in C$, is atomic and live if DAPs are live in $\xi$.

**Proof.** The proof of the claim is similar to the case of algorithm $A_1$ except for the argument in Property P1 Case (d). In case of $A_2$, the $c$.get-data() is executed as a part of $\phi$ and completes before $c$.get-data() is called in $\pi$. By property $C3$ of the data-access primitives, we have $\tau_\phi \leq \tau_\pi$ and hence we cannot have $\pi \preceq \phi$. 

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Expressing an atomic algorithm in terms of the DAP primitives serves multiple purposes. First, describing an algorithm according to templates $A_1$ or $A_2$ allows one to prove that the algorithm is safe (atomic) by just showing that the appropriate DAP properties hold, and the algorithm is live if the implementation of each primitive is live. Secondly, the safety and liveness proofs for more complex algorithms (like ARES in Section 4) become easier as one may reason on the DAP properties that are satisfied by the primitives used, without involving the underlying implementation of those primitives. And last but not least, describing a reconfigurable algorithm using DAPs, provides the flexibility to use different implementation mechanics for the DAPs in each reconfiguration, as long as the DAPs used in a single configuration satisfy the appropriate DAP properties. In Section 4 we discuss how ARES may change the primitives mechanisms in each established configuration without affecting the safety guarantees of the service. Such approaches can adapt to the configuration design, and vary the performance of the service based on the environmental conditions. In other words, ABD [4] can be used for maximum fault tolerance and when majority quorums are used, whereas fast algorithms similar to the ones presented in [9, 3], could be used in configurations that satisfy the appropriate participation bounds.

A.1 Representing Known Algorithms in terms of data-access primitives

Any tag-based algorithm can be transformed into a generic algorithm $A_1$ or $A_2$. A straightforward transformation of any multi-reader multi-writer atomic memory algorithm $A$ is to convert it to $A_1$ by appropriately defining $c.$get-data() action in terms of the protocol for the read operation in $A$; and the $c.$put-data($\langle \tau, v \rangle$) in terms of the write protocol. Since $A$ is an atomic algorithm such implementation of the primitives would satisfy the DAP consistency properties and therefore $A_1$ would also satisfy atomicity. However, such transformation of $A$ is not necessarily efficient in terms of the number of rounds or communication complexity associated with an operation, or even storage cost, as each read and write operation performs both the read and write protocols of the original algorithm $A$.

In this subsection we demonstrate how two well known algorithms for emulating atomic read/write memory can be transformed to their generic equivalent algorithms. In particular, we will present the very celebrated ABD algorithm [4] and the LDR algorithm presented in [10]. For both algorithms we will specify a transformation to a generic algorithm and present the implementations of their data-primitives as well as the primitive handlers.

**MWABD Algorithm.** The multi-writer version of the ABD can be transformed to the generic algorithm $A_1$. Automaton[12] illustrates the three DAP for the ABD algorithm. The get-data primitive encapsulates the query phase of MWABD, while the put-data primitive encapsulates the propagation phase of the algorithm.

Let us now examine if the primitives satisfy properties $C_1$ and $C_2$. We begin with a lemma that shows the monotonicity of the tags at each server.

**Lemma 34.** Let $\sigma$ and $\sigma'$ two states in an execution $\xi$ such that $\sigma$ appears before $\sigma'$ in $\xi$. Then for any server $s \in S$ it must hold that $s.tag|_{\sigma} \leq s.tag|_{\sigma'}$.

**Proof.** According to the algorithm, a server $s$ updates its local tag-value pairs when it receives a message with a higher tag. So if $s.tag|_{\sigma} = \tau$ then in a state $\sigma'$ that appears after $\sigma$ in $\xi$, $s.tag|_{\sigma'} \geq \tau$. 

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Algorithm 12 Implementation of DAP for ABD at each process \( p \) using configuration \( c \)

Let \( \phi \) be a \( \text{c.put-data}(\langle \tau, v \rangle) \) action invoked by \( p_1 \) and \( \gamma \) be a \( \text{c.get-tag()} \) action invoked by \( p_2 \) in a configuration \( c \), such that \( \phi \rightarrow \gamma \) in an execution \( \xi \) of the algorithm. Then \( \gamma \) returns a tag \( \tau_\gamma \geq \tau \).

**Proof.** The lemma follows from the intersection property of quorums. In particular, during the \( \text{c.put-data}(\langle \tau, v \rangle) \) action, \( p_1 \) sends the pair \( \langle \tau, v \rangle \) to all the servers in \( c.Servers \) and waits until all the servers in a quorum \( Q_i \in c.Quorums \) reply. When those replies are received then the action completes.

During a \( \text{c.get-data()} \) action on the other hand, \( p_2 \) sends query messages to all the servers in \( c.Servers \) and waits until all servers in a quorum \( Q_j \in c.Quorums \) (not necessarily different than \( Q_i \)) reply. By definition \( Q_i \cap Q_j \neq \emptyset \), thus any server \( s \in Q_i \cap Q_j \) reply to both \( \phi \) and \( \gamma \) actions. By Lemma 34 and since \( s \) received a tag \( \tau \), then \( s \) replies to \( p_2 \) with a tag \( \tau_s \geq \tau \). Since \( \gamma \) returns the maximum tag it discovers then \( \tau_\gamma \geq \tau_s \). Therefore \( \tau_\gamma \geq \tau \) and this completes the proof.

With similar arguments and given that each value is associated with a unique tag then we can show the following lemma.

**Lemma 36.** Let \( \pi \) be a \( \text{c.put-data}(\langle \tau, v \rangle) \) action invoked by \( p_1 \) and \( \phi \) be a \( \text{c.get-data()} \) action invoked by \( p_2 \) in a configuration \( c \), such that \( \pi \rightarrow \phi \) in an execution \( \xi \) of the algorithm. Then \( \phi \) returns a tag-value \( \langle \tau_\phi, v_\phi \rangle \) such that \( \tau_\phi \geq \tau \).

Finally we can now show that property \( \text{C2} \) also holds.
Lemma 37. If φ is a c.get-data() that returns \(\langle \tau_\pi, v_\pi \rangle \in \mathcal{T} \times \mathcal{V}\), then there exists \(\pi\) such that \(\pi\) is a c.put-data(\(\langle \tau_\pi, v_\pi \rangle\)) and \(\phi \not\rightarrow \pi\).

Proof. This follows from the facts that (i) servers set their tag-value pair to a pair received by a put-data action, and (ii) a get-data action returns a tag-value pair that it received from a server. So if a c.get-data(\(\phi\)) operation \(\phi\) returns a tag-value pair \(\langle \tau_\pi, v_\pi \rangle\), there should be a server \(s\) that replied to that operation with \(\langle \tau_\pi, v_\pi \rangle\), and \(s\) received \(\langle \tau_\pi, v_\pi \rangle\) from some c.put-data(\(\langle \tau_\pi, v_\pi \rangle\)) action, \(\pi\). Thus, \(\pi\) can proceed or be concurrent with \(\phi\), and hence \(\phi \not\rightarrow \pi\).

LDR Algorithm. The LDR algorithm [10] was designed with large objects in mind, and for that reason it decouples the meta-data associated with each atomic object from the actual object value. So the algorithm considers the existence of two sets of servers: (a) the directory servers that maintain the information about the latest tags in the system, and (b) the replica servers that maintain the replica values (also associated with particular tags). So a configuration \(c\) must include two sets of servers: Directories(\(c\)) \(\subset\) c_servers and Replicas(\(c\)) \(\subset\) c_servers. Automaton [13] shows the specification along with the handlers of each DAP. With this specification LDR can be transformed to the template of algorithm \(A_2\) where essentially the c.get-data() primitive encapsulates the specification of the read operation as defined in LDR.

B TREAS Correctness and Liveness

Lemma. [5] The data-access primitives, i.e., get-tag, get-data and put-data primitives, implemented in the TREAS algorithm satisfy the consistency properties.

Proof. As mentioned above we are concerned with only configuration \(c\), and therefore, in our proofs we will be concerned with only one configuration. Let \(\alpha\) be some execution of TREAS, then we consider two cases for \(\pi\) for proving property C1: \(\pi\) is a get-tag operation, or \(\pi\) is a get-data primitive.

Case (a): \(\phi\) is c.put-data(\(\langle \tau_\phi, v_\phi \rangle\)) and \(\pi\) is a c.get-tag() returns \(\tau_\pi \in \mathcal{T}\). Let \(c_\phi\) and \(c_\pi\) denote the clients that invokes \(\phi\) and \(\pi\) in \(\alpha\). Let \(S_\phi \subset S\) denote the set of \(\lceil \frac{n+k}{2} \rceil\) servers that responds to \(c_\phi\), during \(\phi\). Denote by \(S_\pi\) the set of \(\lceil \frac{n+k}{2} \rceil\) servers that responds to \(c_\pi\), during \(\pi\). Let \(T_1\) be a point in execution \(\alpha\) after the completion of \(\phi\) and before the invocation of \(\pi\). Because \(\pi\) is invoked after \(T_1\), therefore, at \(T_1\) each of the servers in \(S_\phi\) contains \(t_\phi\) in its List variable. Note that, once a tag is added to List, it is never removed. Therefore, during \(\pi\), any server in \(S_\phi \cap S_\pi\) responds with List containing \(t_\phi\) to \(c_\pi\). Note that since \(|S_\sigma^*| = |S_\pi| = \lceil \frac{n+k}{2} \rceil\) implies \(|S_\sigma^* \cap S_\pi| \geq k\), and hence \(t_{\max}^{dec}\) at \(c_\pi\), during \(\pi\) is at least as large as \(t_\phi\), i.e., \(t_\pi \geq t_\phi\). Therefore, it suffices to prove our claim with respect to the tags and the decodability of its corresponding value.

Case (b): \(\phi\) is c.put-data(\(\langle \tau_\phi, v_\phi \rangle\)) and \(\pi\) is a c.get-data() returns \(\langle \tau_\pi, v_\pi \rangle \in \mathcal{T} \times \mathcal{V}\). As above, let \(c_\phi\) and \(c_\pi\) be the clients that invokes \(\phi\) and \(\pi\). Let \(S_\phi\) and \(S_\pi\) be the set of servers that responds to \(c_\phi\) and \(c_\pi\), respectively. Arguing as above, \(|S_\sigma^* \cap S_\pi| \geq k\) and every server in \(S_\phi \cap S_\pi\) sends \(t_\phi\) in response to \(c_\phi\), during \(\pi\), in their List's and hence \(t_\phi \in Tags_{\max}^{\pi\max}\). Now, because \(\pi\) completes in \(\alpha\), hence we have \(t_{\max}^* = t_{\max}^{dec}\). Note that \(|\max Tags_{\max}^{\pi}\) \(\geq k\) and \(\max Tags_{\max}^{\pi}\) so \(t_\pi \geq \max Tags_{\max}^{\pi}\) \(\geq \max Tags_{\max}^{\phi}\) \(\geq t_\phi\). Note that each tag is always associated with its corresponding value \(v_\pi\), or the corresponding coded elements \(\Phi_s(v_\pi)\) for \(s \in S\).
Algorithm 13 Implementation of DAP for LDR at process $p$ using configuration $c$

Data-Access Primitives at each process $p$:

procedure c.get-tag()
  send (QUERY-TAG-LOCATION) to each $s \in Dir(c)$
  until $\exists Q, Q \in Majority(Dir(c))$ s.t.
    $p$ receives $(\tau_s, loc_s)$ from $\forall s \in Q$
  $\tau_{max} \leftarrow \max\{\{\tau_s : p, p \in Q \text{ received } (\tau_s, loc_s) \text{ from } s\}\}$
  return $\tau_{max}$
end procedure

procedure c.put-data($\tau$, $v$))
  send (PUT-DATA, $(\tau, v)$) to $2f + 1$ servers in Rep(c)
  until $p$ receives ACK from a set $U$ of $f + 1$ servers in Rep(c)
  send (PUT-METADATA, $(\tau, U)$) to all servers in Dir(c)
  until $p$ receives ACK from a majority servers in Dir(c)
end procedure

Primitive Handlers at each Directory server $s \in Dir(c)$:
Upon receive (QUERY-TAG-LOCATION) from $q$
  send $(\tau, loc)$ to $q$
end receive

Upon receive (PUT-METADATA, $(\tau_{in}, loc_{in})$) from $q$
if $\tau_{in} > \tau$ then
  $(\tau, loc) \leftarrow (\tau_{in}, loc_{in})$
  send ACK to $q$
end receive

Primitive Handlers at each Replica server $s \in Rep(c)$:
Upon receive (GET-DATA) from $q$
  send $(\tau, v)$ to $q$
end receive

Upon receive (PUT-DATA, $(\tau_{in}, v_{in})$) from $q$
if $\tau_{in} > \tau$ then
  $(\tau, v) \leftarrow (\tau_{in}, v_{in})$
  send ACK to $q$
end receive

Next, we prove the C2 property of DAP for the TREAS algorithm. Note that the initial values of the List variable in each server $s$ in $S$ is $\{\{(t_0, \Phi_s(v_s))\}\}$. Moreover, from an inspection of the steps of the algorithm, new tags in the List variable of any servers of any servers is introduced via put-data operation. Since $t_\pi$ is returned by a get-tag or get-data operation then it must be that either $t_\pi = t_0$ or $t_\pi > t_0$. In the case where $t_\pi = t_0$ then we have nothing to prove. If $t_\pi > t_0$ then there must be a put-data($t_\pi$, $v_\pi$) operation $\phi$. To show that for every $\pi$ it cannot be that $\phi$ completes before $\pi$, we adopt by a contradiction. Suppose for every $\pi$, $\phi$ completes before $\pi$ begins, then clearly $t_\pi$ cannot be returned $\phi$, a contradiction. \hfill $\Box$

**Theorem.** 9 [Liveness] Let $\beta$ denote a well-formed execution of TREAS, with $[n, k]$, where $n$ is the number of servers and $k > n/3$, and $\delta$ be the maximum number of write operations concurrent with any valid read operation then the read and write operations $\beta$ are live.

**Proof.** Note that in the read and write operation the get-tag and put-data operations initiated by any non-faulty client always complete. Therefore, the liveness property with respect to any write operation is clear because it uses only get-tag and get-data operations of the DAP. So, we focus on proving the liveness property of any read operation $\pi$, specifically, the get-data operation completes.
Let $\alpha$ be and execution of TREAS and let $c_{\alpha^*}$ and $c_{\pi}$ be the clients that invokes the write operation $\sigma^*$ and read operation $c_{\sigma^*}$, respectively.

Let $S_{\sigma^*}$ be the set of $\left\lceil \frac{n+k}{2} \right\rceil$ servers that responds to $c_{\sigma^*}$, in the put-data operations, in $\sigma^*$. Let $S_{\pi^*}$ be the set of $\left\lceil \frac{n+k}{2} \right\rceil$ servers that responds to $c_{\pi}$ during the get-data step of $\pi$. Note that in $\alpha$ at the point execution $T_1$, just before the execution of $\pi$, none of the the write operations in $\Lambda$ is complete. Observe that, by algorithm design, the coded-elements corresponding to $n$ we have the worst case communication cost of a write operation is put-data variables, which are metadata. However, in the most Lemma 40. The communication cost associated with a successful read operation in TREAS is at most $\delta + \frac{n}{k}$.

Proof. During read operation, in the most Lemma 40. The communication cost associated with a successful read operation in TREAS is at most $\delta + \frac{n}{k}$.

Proof. During read operation, the reader sends each server the coded-elements corresponding to tag $t \in S_{\sigma^*}$, in the get-tag phase the servers responds with their highest tags variables, which are metadata. However, in the put-data phase, the reader sends each server the coded elements of size $\frac{k}{\delta}$ each, and hence the total cost of communication for this is $\delta + \frac{n}{k}$. Therefore, we have the worst case communication cost of a write operation is $\delta + \frac{n}{k}$.
Proof. During read operation, in the get-data phase the servers responds with their \textit{List} variables and hence each such list is of size at most \((\delta + 1)\frac{1}{k}\), and then counting all such responses give us \((\delta + 1)\frac{n}{k}\). In the put-data phase, the reader sends each server the coded elements of size \(\frac{1}{k}\) each, and hence the total cost of communication for this is \(\frac{n}{k}\). Therefore, we have the worst case communication cost of a read operation is \((\delta + 2)\frac{n}{k}\).

From the above Lemmas we get.

\textbf{Theorem. 3} The \textsc{Treas} algorithm has: (i) storage cost \((\delta + 1)\frac{n}{k}\), (ii) communication cost for each write at most to \(\frac{n}{k}\), and (iii) communication cost for each read at most to \((\delta + 2)\frac{n}{k}\).

C \hspace{1cm} \textbf{ARES Safety}

\textbf{Notations and definitions.} For a server \(s\), we use the notation \(s.\text{var}|_{\sigma}\) to refer to the value of the state variable \(\text{var}\), in \(s\), at a state \(\sigma\) of an execution \(\xi\). If server \(s\) crashes at a state \(\sigma_f\) in an execution \(\xi\) then \(s.\text{var}|_{\sigma} \triangleq s.\text{var}|_{\sigma_f}\) for any state variable \(\text{var}\) and for any state \(\sigma\) that appears after \(\sigma_f\) in \(\xi\).

\textbf{Consensus instance in a configuration} We assume that the servers in each configuration \(c\) implements a consensus service \(c.\text{Con}\), where any client is allowed to proposed values from the set \(C\), and \(c.\text{Con}\) satisfies the following properties

\textbf{Definition 41.} For a configuration \(c \in C\), \(c.\text{Con}\) must satisfy the following properties:

\textit{Agreement:} No two processes that participate in \(c.\text{Con}\) decide a different value.

\textit{Validity:} If any process decides a value \(c'\) then \(c'\) was proposed by some process.

\textit{Termination:} Every correct process decides.

\textbf{Definition 42} (Tag of a configuration). Let \(c \in C\) be a configuration, \(\sigma\) be a state in some execution \(\xi\) then we define the tag of \(c\) at state \(\sigma\) as \(\text{tag}(c)|_{\sigma} \triangleq \min_{Q \in c.\text{Quorums}} \max_{s \in Q} (s.\text{tag}|_{\sigma})\). We often drop the suffix \(\igma\), and simply denote as \(\text{tag}(c)\), when the state in the execution is clear from the context.

\textbf{Definition 43.} Let \(c^p_{\sigma} = p.\text{cseq}|_{\sigma}\). Then we define as \(\mu(c^p_{\sigma}) \triangleq \max\{i : c^p_{\sigma}[i].\text{status} = F\}\) and \(\nu(c^p_{\sigma}) \triangleq |c^p_{\sigma}|\), where \(|c^p_{\sigma}|\) is the number of elements in configuration vector \(c^p_{\sigma}\) that are not equal to \(\perp\).

\textbf{Definition 44} (Prefix order). Let \(x\) and \(y\) be any two configuration sequences. We say that \(x\) is a prefix of \(y\), denoted by \(x \preceq_p y\), if \(x[j].\text{cfg} = y[j].\text{cfg}\), for all \(j\) such that \(x[j] \neq \perp\).

Next we analyze the properties that we can achieve through our reconfiguration algorithm. The first lemma shows that any two configuration sequences have the same configuration identifiers in the same indexes.

\textbf{Lemma 45.} For any reconfigurer \(r\) that invokes an \textit{reconfig}(c) action in an execution \(\xi\) of the algorithm, If \(r\) chooses to install \(c\) in index \(k\) of its local \(r.\text{cseq}\) vector, then \(r\) invokes the \(\text{Cons}[k - 1].\text{propose}(c)\) instance over configuration \(r.\text{cseq}[k - 1].\text{cfg}\).
Proof. It follows directly from the algorithm. □

Lemma 46. If a server $s$ sets $s.nextC$ to $\langle c, F \rangle$ at some state $\sigma$ in an execution $\xi$ of the algorithm, then $s.nextC = \langle c, F \rangle$ for any state $\sigma'$ that appears after $\sigma$ in $\xi$.

Proof. Notice that a server $s$ updates the $s.nextC$ variable for some specific configuration $c_k$ in a state $st$ if: (i) $s$ did not receive any value for $c_k$ before (and thus $nextC = \bot$), or (ii) $s$ received a tuple $\langle c, P \rangle$ and before $\sigma$ received the tuple $\langle c', F \rangle$. By Observation 44 $c = c'$ as $s$ updates the $s.nextC$ of the same configuration $c_k$. Once the tuple becomes equal to $\langle c, F \rangle$ then $s$ does not satisfy the update condition for $c_k$, and hence in any state $\sigma'$ after $\sigma$ it does not change $\langle c, F \rangle$. □

Lemma 47 (Configuration Uniqueness). For any processes $p, q \in \mathcal{I}$ and any states $\sigma_1, \sigma_2$ in an execution $\xi$, it must hold that $c^p[i].cfg = c^q[i].cfg, \forall i s.t. c^p[i].cfg, c^q[i].cfg \neq \bot$.

Proof. The lemma holds trivially for $c^p[0].cfg = c^q[0].cfg = c_0$. So in the rest of the proof we focus in the case where $i > 0$. Let us assume w.l.o.g. that $\sigma_1$ appears before $\sigma_2$ in $\xi$.

According to our algorithm a process $p$ sets $p.cseq[i].cfg$ to a configuration identifier $c$ in two cases: (i) either it received $c$ as the result of the consensus instance in configuration $p.cseq[i−1].cfg$, or (ii) $p$ receives $s.nextC.cfg = c$ from a server $s \in p.cseq[i−1].cfg.Servers$. Note here that (i) is possible only when $p$ is a reconfigurer and attempts to install a new configuration. On the other hand (ii) may be executed by any process in any operation that invokes the read-config action. We are going to proof this lemma by induction on the configuration index.

Base case: The base case of the lemma is when $i = 1$. Let us first assume that $p$ and $q$ receive $c_p$ and $c_q$, as the result of the consensus instance at $p.cseq[0].cfg$ and $q.cseq[0].cfg$ respectively. By Lemma 45, since both processes want to install a configuration in $i = 1$, then they have to run $Cons[0]$ instance over the configuration stored in their local $cseq[0].cfg$ variable. Since $p.cseq[0].cfg = q.cseq[0].cfg = c_0$ then both $Cons[0]$ instances run over the same configuration $c_0$ and according to Observation 44 return the same value, say $c_1$. Hence $c_p = c_q = c_1$ and $p.cseq[1].cfg = q.cseq[1].cfg = c_1$.

Let us examine the case now where $p$ or $q$ assign a configuration $c$ they received from some server $s \in c_0.Servers$. According to the algorithm only the configuration that has been decided by the consensus instance on $c_0$ is propagated to the servers in $c_0.Servers$. If $c_1$ is the decided configuration, then $\forall s \in c_0.Servers$ such that $s.nextC(c_0) \neq \bot$, it holds that $s.nextC(c_0) = \langle c_1, * \rangle$. So if $p$ or $q$ set $p.cseq[1].cfg$ or $q.cseq[1].cfg$ to some received configuration, then $p.cseq[1].cfg = q.cseq[1].cfg = c_1$ in this case as well.

Hypothesis: We assume that $c^p[k][i] = c^q[k][i]$ for some $k, k \geq 1$.

Induction Step: We need to show that the lemma holds for $i = k + 1$. If both processes retrieve $p.cseq[k+1].cfg$ and $q.cseq[k+1].cfg$ through consensus, then both $p$ and $q$ run consensus over the previous configuration. Since according to our hypothesis $c^p[k][i] = c^q[k][i]$ then both processes will receive the same decided value, say $c_{k+1}$, and hence $p.cseq[k+1].cfg = q.cseq[k+1].cfg = c_{k+1}$. Similar to the base case, a server in $c_k.Servers$ only receives the configuration $c_{k+1}$ decided by the consensus instance run over $c_k$. So processes $p$ and $q$ can only receive $c_{k+1}$ from some server in $c_k.Servers$ so they can only assign $p.cseq[k+1].cfg = q.cseq[k+1].cfg = c_{k+1}$ at Line A57. That completes the proof.

Lemma 47 showed that any two operations store the same configuration in any cell $k$ of their $cseq$ variable. It is not known however if the two processes discover the same number of configuration
ids. In the following lemmas we will show that if a process learns about a configuration in a cell \(c\) then it also learns about all configuration ids for every index \(i\), such that \(0 \leq i \leq k - 1\).

**Lemma 48.** In any execution \(\xi\) of the algorithm, for any process \(p \in I\), \(c^p_\sigma[i] \neq \bot\) in some state \(\sigma\) in \(\xi\), then \(c^p_\sigma[i] \neq \bot\) in any state \(\sigma'\) that appears after \(\sigma\) in \(\xi\).

**Proof.** A value is assigned to \(c^p_\sigma[i]\) either after the invocation of a consensus instance, or while executing the read-config action. Since any configuration proposed for installation cannot be \(\bot\) (A57), and since there is at least one configuration proposed in the consensus instance (the one from \(p\)), then by the validity of the consensus service the decision will be a configuration \(c \neq \bot\). Thus, in this case \(c^p_\sigma[i]\) cannot be \(\bot\). Also in the read-config procedure, \(c^p_\sigma[i]\) is assigned to a value different than \(\bot\) according to Line A517. Hence, if \(c^p_\sigma[i] \neq \bot\) at state \(\sigma\) then it cannot become \(\bot\) in any state \(\sigma'\) after \(\sigma\) in execution \(\xi\).

**Lemma 49.** Let \(\sigma_1\) be some state in an execution \(\xi\) of the algorithm. Then for any process \(p\), if \(k = \max\{i : c^p_{\sigma_1}[i] \neq \bot\}\), then \(c^p_{\sigma_1}[j] \neq \bot\), for \(0 \leq j < k\).

**Proof.** Let us assume to derive contradiction that there exists \(j < k\) such that \(c^p_{\sigma_1}[j] = \bot\) and \(c^p_{\sigma_1}[j + 1] \neq \bot\). Suppose w.l.o.g. that \(j = k - 1\) and that \(\sigma_1\) is the state immediately after the assignment of a value to \(c^p_{\sigma_1}[k]\), say \(c_k\). Since \(c^p_{\sigma_1}[k] \neq \bot\), then \(p\) assigned \(c_k\) to \(c^p_{\sigma_1}[k]\) in one of the following cases: (i) \(c_k\) was the result of the consensus instance, or (ii) \(p\) received \(c_k\) from a server during a read-config action. The first case is trivially impossible as according to Lemma 45 \(p\) decides for \(k\) when it runs consensus over configuration \(c^p_{\sigma_1}[k - 1]\). \(cfg\). Since this is equal to \(\bot\), then we cannot run consensus over a non-existent set of processes. In the second case, \(p\) assigns \(c^p_{\sigma_1}[k] = c_k\). The value \(c_k\) was however obtained when \(p\) invoked get-next-config on configuration \(c^p_{\sigma_1}[k - 1]\). \(cfg\). In that action, \(p\) sends read-config messages to the servers in \(c^p_{\sigma_1}[k - 1]\). \(Servers\) and waits until a quorum of servers replies. Since we assigned \(c^p_{\sigma_1}[k] = c_k\) it means that get-next-config terminated at some state \(\sigma'\) before \(\sigma_1\) in \(\xi\), and thus: (a) a quorum of servers in \(c^p_{\sigma_1}[k - 1]\). \(cfg\). \(Servers\) replied, and (b) there exists a server \(s\) among those that replied with \(c_k\). According to our assumption however, \(c^p_{\sigma_1}[k - 1] = \bot\) at \(\sigma_1\). So if state \(\sigma'\) is before \(\sigma_1\) in \(\xi\), then by Lemma 48 it follows that \(c^p_{\sigma_1}[k - 1] = \bot\). This however implies that \(p\) communicated with an empty configuration, and thus no server replied to \(p\). This however contradicts the assumption that a server replied with \(c_k\) to \(p\).

Since any process traverses the configuration sequence starting from the initial configuration \(c_0\), then with a simple induction we can show that \(c^p_{\sigma_1}[j] \neq \bot\), for \(0 \leq j \leq k\).

We can now move to an important lemma that shows that any read-config action returns an extension of the configuration sequence returned by any previous read-config action. First, we show that the last finalized configuration observed by any read-config action is at least as recent as the finalized configuration observed by any subsequent read-config action.

**Lemma 50.** If at a state \(\sigma\) of an execution \(\xi\) of the algorithm, if \(\mu(c^p_\sigma) = k\) for some process \(p\), then for any element \(0 \leq j < k\), \(\exists Q \in c^p_\sigma[j].cfg.\) Quorums such that \(\forall s \in Q, s.nextC(c^p_\sigma[j].cfg) = c^p_\sigma[j + 1]\).

**Proof.** This lemma follows directly from the algorithm. Notice that whenever a process assigns a value to an element of its local configuration (Lines A49 and A517), it then propagates this value
to a quorum of the previous configuration (Lines A4[18] and A5[18]). So if a process \( p \) assigned \( c_j \) to an element \( c_{\sigma'}[j] \) in some state \( \sigma' \) in \( \xi \), then \( p \) may assign a value to the \( j + 1 \) element of \( c_{\sigma'}[j + 1] \) only after put-config\( (c_{\sigma'}[j - 1].cfg, c_{\sigma'}[j]) \) occurs. During put-config action, \( p \) propagates \( c_{\sigma'}[j] \) in a quorum \( Q \in c_{\sigma'}[j - 1].cfg.Q\)orums. Hence, if \( c_{\sigma'}[k] \neq \bot \), then \( p \) propagated each \( c_{\sigma'}[j] \), for \( 0 < j \leq k \) to a quorum of servers \( Q \in c_{\sigma'}[j - 1].cfg.Q\)orums. And this completes the proof.

**Lemma 51** (Configuration Prefix). Let \( \pi_1 \) and \( \pi_2 \) two completed read-config actions invoked by processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \) in an execution \( \xi \). Let \( \sigma_1 \) be the state after the response step of \( \pi_1 \) and \( \sigma_2 \) the state after the response step of \( \pi_2 \). Then \( c_{\sigma_1} \preceq c_{\sigma_2} \).

**Proof.** Let \( \nu_1 = \nu(c_{\sigma_1}) \) and \( \nu_2 = \nu(c_{\sigma_2}) \). By Lemma 47 for any \( i \) such that \( c_{\sigma_1}[i] \neq \bot \) and \( c_{\sigma_2}[i] \neq \bot \), then \( c_{\sigma_1}[i].cfg = c_{\sigma_2}[i].cfg \). Also from Lemma 49 we know that for \( 0 \leq j \leq \nu_1, c_{\sigma_1}[j] \neq \bot \), and \( 0 \leq j \leq \nu_2, c_{\sigma_2}[j] \neq \bot \). So if we can show that \( \nu_1 \leq \nu_2 \) then the lemma follows.

Let \( \mu = \mu(c_{\sigma_2}) \) be the last finalized element which \( p_2 \) established in the beginning of the read-config action \( \pi_2 \) (Line A5[2]) at some state \( \sigma' \) before \( \sigma_2 \). It is easy to see that \( \mu \leq \nu_2 \). If \( \nu_1 \leq \mu \) then \( \mu \leq \nu_2 \) and the lemma follows. Thus, it remains to examine the case where \( \mu < \nu_1 \). Notice that since \( \pi_1 \rightarrow \pi_2 \) then \( \sigma_1 \) appears before \( \sigma' \) in execution \( \xi \). By Lemma 50, we know that by \( \sigma_1 \), \( \exists Q \in c_{\sigma_1}[j].cfg.Q\)orums, for \( 0 \leq j < \nu_1 \), such that \( \forall s \in Q, s.nextC = c_{\sigma_1}[j + 1] \). Since \( \mu < \nu_1 \), then it must be the case that \( \exists Q \in c_{\sigma_1}[\mu].cfg.Q\)orums such that \( \forall s \in Q, s.nextC = c_{\sigma_1}[\mu + 1] \). But by Lemma 47 we know that \( c_{\sigma_1}[\mu].cfg = c_{\sigma_2}[\mu].cfg \). Let \( Q' \) be the quorum that replies to the read-next-config occurred in \( p_2 \), on configuration \( c_{\sigma_2}[\mu].cfg \). By definition \( Q \cap Q' = \emptyset \), thus there is a server \( s \in Q \cap Q' \) that sends \( s.nextC = c_{\sigma_1}[\mu + 1] \) to \( p_2 \) during \( \pi_2 \). Since \( c_{\sigma_1}[\mu + 1] \neq \bot \) then \( p_2 \) assigns \( c_{\sigma_2}[\mu + 1] = c_{\sigma_1}[\mu + 1] \), and repeats the process in the configuration \( c_{\sigma_2}[\mu + 1].cfg \). Since every configuration \( c_{\sigma_2}[\mu].cfg \), for \( \mu \leq j < \nu_1 \), has a quorum of servers with \( s.nextC \), then by a simple induction it can be shown that the process will be repeated for at least \( \nu_1 - \mu \) iterations, and every configuration \( c_{\sigma_2}[\mu].cfg \), for \( \mu \leq j \leq \nu_1 \), has a quorum of servers \( s.nextC \) at some state \( \sigma'' \) before \( \sigma_2 \). Thus, \( c_{\sigma_2}[\mu] = c_{\sigma_2}[\mu] \), for \( 0 \leq j \leq \nu_1 \). Hence \( \nu_1 \leq \nu_2 \) and the lemma follows in this case as well.

Thus far we focused on the configuration member of each element in \( cseq \). As operations do get in account the status of a configuration, i.e. \( P \) or \( F \), in the next lemma we will examine the relationship of the last finalized configuration as detected by two operations. First we present a lemma that shows the monotonicity of the finalized configurations.

**Lemma 52.** Let \( \sigma \) and \( \sigma' \) two states in an execution \( \xi \) such that \( \sigma \) appears before \( \sigma' \) in \( \xi \). Then for any process \( p \) must hold that \( \mu(c_{\sigma}) \leq \mu(c_{\sigma'}) \).

**Proof.** This lemma follows from the fact that if a configuration \( k \) is such that \( c_{\sigma}[k].status = F \) at a state \( \sigma \), then \( p \) will start any future read-config action from a configuration \( c_{\sigma'}[j].cfg \) such that \( j \geq k \). But \( c_{\sigma'}[j].cfg \) is the last finalized configuration at \( \sigma' \) and hence \( \mu(c_{\sigma'}) \geq \mu(c_{\sigma}) \).

**Lemma 53** (Configuration Progress). Let \( \pi_1 \) and \( \pi_2 \) two completed read-config actions invoked by processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \) in an execution \( \xi \). Let \( \sigma_1 \) be the state after the response step of \( \pi_1 \) and \( \sigma_2 \) the state after the response step of \( \pi_2 \). Then \( \mu(c_{\sigma_1}) \leq \mu(c_{\sigma_2}) \).

**Proof.** By Lemma 51 it follows that \( c_{\sigma_1} \) is a prefix of \( c_{\sigma_2} \). Thus, if \( \nu_1 = \nu(c_{\sigma_1}) \) and \( \nu_2 = \nu(c_{\sigma_2}) \), \( \nu_1 \leq \nu_2 \). Let \( \mu_1 = \mu(c_{\sigma_1}) \), such that \( \mu_1 \leq \nu_1 \), be the last element in \( c_{\sigma_1} \), where \( c_{\sigma_1}[\mu].status = F \). Let now \( \mu_2 = \mu(c_{\sigma_2}) \), be the last element which \( p_2 \) obtained in Line A4[2] during \( \pi_2 \) such that
\[c^\sigma_{s_1}^{p_1}[\mu_2].status = F\] in some state \(\sigma'\) before \(\sigma_2\). If \(\mu_2 \geq \mu_1\), and since \(\sigma_2\) is after \(\sigma'\), then by Lemma 52 \(\mu_2 \leq \mu(c^\sigma_{s_2})\) and hence \(\mu_1 \leq \mu(c^\sigma_{s_2})\) as well.

It remains to examine the case where \(\mu_2 < \mu_1\). Process \(p_1\) sets the status of \(c^\sigma_{s_1}[\mu_1]\) to \(F\) in two cases: (i) either when finalizing a reconfiguration, or (ii) when receiving an \(s.nextC = \langle c^\sigma_{s_1}[\mu_1].cfg, F \rangle\) from some server \(s\) during a read-config action. In both cases \(p_1\) propagates the \(\langle c^\sigma_{s_1}[\mu_1].cfg, F \rangle\) to a quorum of servers in \(c^\sigma_{s_2}[\mu_1 - 1].cfg\) before completing. We know by Lemma 51 that since \(\pi_1 \rightarrow \pi_2\) then \(c^\sigma_{s_1}\) is a prefix in terms of configurations of the \(c^\sigma_{s_2}\). So it must be the case that \(\mu_2 < \mu_1 \leq \nu(c^\sigma_{s_2})\). Thus, during \(\pi_2\), \(p_2\) starts from the configuration at index \(\mu_2\) and in some iteration performs get-next-config in configuration \(c^\sigma_{s_2}[\mu_1 - 1]\). According to Lemma 47, \(c^\sigma_{s_1}[\mu_1 - 1].cfg = c^\sigma_{s_2}[\mu_1 - 1].cfg\). Since \(\pi_1\) completed before \(\pi_2\), then it must be the case that \(\sigma_1\) appears before \(\sigma'\) in \(\xi\). However, \(p_2\) invokes the get-next-config operation in a state \(\sigma''\) which is either equal to \(\sigma'\) or appears after \(\sigma'\) in \(\xi\). Thus, \(\sigma''\) must appear after \(\sigma_1\) in \(\xi\). From that it follows that when the get-next-config is executed by \(p_2\) there is already a quorum of servers in \(c^\sigma_{s_2}[\mu_1 - 1].cfg\), say \(Q_1\), that received \(\langle c^\sigma_{s_1}[\mu_1].cfg, F \rangle\) from \(p_1\). Since, \(p_2\) waits from replies from a quorum of servers from the same configuration, say \(Q_2\), and since the nextC variable at each server is monotonic (Lemma 46), then there is a server \(s \in Q_1 \cap Q_2\), such that \(s\) replies to \(p_2\) with \(s.nextC = \langle c^\sigma_{s_1}[\mu_1].cfg, F \rangle\). So, \(c^\sigma_{s_2}[\mu_1]\) gets \(\langle c^\sigma_{s_1}[\mu_1].cfg, F \rangle\), and hence \(\mu(c^\sigma_{s_2}) \geq \mu_1\) in this case as well. This completes our proof.

**Theorem 54.** Let \(\pi_1\) and \(\pi_2\) two completed read-config actions invoked by processes \(p_1, p_2 \in \mathcal{I}\) respectively, such that \(\pi_1 \rightarrow \pi_2\) in an execution \(\xi\). Let \(\sigma_1\) be the state after the response step of \(\pi_1\) and \(\sigma_2\) the state after the response step of \(\pi_2\). Then the following properties hold:

1. \(c^\sigma_{s_2}[i].cfg = c^\sigma_{s_1}[i].cfg\) for \(1 \leq i \leq \nu(c^\sigma_{s_1})\),
2. \(c^\sigma_{s_1} \preceq_p c^\sigma_{s_2}\), and
3. \(\mu(c^\sigma_{s_1}) \leq \mu(c^\sigma_{s_2})\)

**Proof.** Statements (a), (b) and (c) follow from Lemmas 47, 51, and 52.

**Lemma.** 45 For any reconfigurer \(r\) that invokes an reconfig(c) action in an execution \(\xi\) of the algorithm, If \(r\) chooses to install \(c\) in index \(k\) of its local r.cseq vector, then \(r\) invokes the Cons\([k - 1].propose(c)\) instance over configuration \(r.cseq[k - 1].cfg\).

**Proof.** It follows directly from the algorithm.

**Lemma.** 46 If a server \(s\) sets \(s.nextC\) to \(\langle c, F \rangle\) at some state \(\sigma\) in an execution \(\xi\) of the algorithm, then \(s.nextC = \langle c, F \rangle\) at any state \(\sigma'\) that appears after \(\sigma\) in \(\xi\).

**Proof.** Notice that a server \(s\) updates the \(s.nextC\) variable for some specific configuration \(c_k\) in a state \(st\) if: (i) \(s\) did not receive any value for \(c_k\) before (and thus \(nextC = \bot\)), or (ii) \(s\) received a tuple \(\langle c, P \rangle\) and before \(\sigma\) received the tuple \(\langle c', F \rangle\). By Observation 41, \(c = c'\) as \(s\) updates the \(s.nextC\) of the same configuration \(c_k\). Once the tuple becomes equal to \(\langle c, F \rangle\) then \(s\) does not satisfy the update condition for \(c_k\), and hence in any state \(\sigma'\) after \(\sigma\) it does not change \(\langle c, F \rangle\).

**Lemma.** 47 [Configuration Uniqueness] For any processes \(p, q \in \mathcal{I}\) and any states \(\sigma_1, \sigma_2\) in an execution \(\xi\), it must hold that \(c^\sigma_{p}[i].cfg = c^\sigma_{q}[i].cfg\), \(\forall i\) s.t. \(c^\sigma_{p}[i].cfg, c^\sigma_{q}[i].cfg \neq \bot\).
Proof. The lemma holds trivially for \( c_{p_1}^0[0].cfg = c_{q_2}^0[0].cfg = c_0 \). So in the rest of the proof we focus in the case where \( i > 0 \). Let us assume w.l.o.g. that \( \sigma_1 \) appears before \( \sigma_2 \) in \( \xi \).

According to our algorithm a process \( p \) sets \( p.cseq[i].cfg \) to a configuration identifier \( c \) in two cases: (i) either it received \( c \) as the result of the consensus instance in configuration \( p.cseq[i−1].cfg \), or (ii) \( p \) receives \( s.nextC.cfg = c \) from a server \( s \in p.cseq[i−1].cfg\).Servers. Note here that (i) is possible only when \( p \) is a reconfigurer and attempts to install a new configuration. On the other hand (ii) may be executed by any process in any operation that invokes the read-config action. We are going to proof this lemma by induction on the configuration index.

Base case: The base case of the lemma is when \( i = 1 \). Let us first assume that \( p \) and \( q \) receive \( c_p \) and \( c_q \), as the result of the consensus instance at \( p.cseq[0].cfg \) and \( q.cseq[0].cfg \) respectively. By Lemma 45 since both processes want to install a configuration in \( i = 1 \), then they have to run \( Cons[0] \) instance over the configuration stored in their local \( cseq[0].cfg \) variable. Since \( p.cseq[0].cfg = q.cseq[0].cfg = c_0 \) then both \( Cons[0] \) instances run over the same configuration \( c_0 \) and according to Definition 41 return the same value, say \( c_1 \). Hence \( p.cseq[1].cfg = q.cseq[1].cfg = c_1 \).

Let us examine the case now where \( p \) or \( q \) assign a configuration \( c \) they received from some server \( s \in c_0.Servers \). According to the algorithm only the configuration that has been decided by the consensus instance on \( c_0 \) is propagated to the servers in \( c_0.Servers \). If \( c_1 \) is the decided configuration, then \( \forall s \in c_0.Servers \) such that \( s.nextC(c_0) \neq \bot \), it holds that \( s.nextC(C_0) = (c_1,*) \). So if \( p \) or \( q \) set \( p.cseq[1].cfg = q.cseq[1].cfg \) to some received configuration, then \( p.cseq[1].cfg = q.cseq[1].cfg = c_1 \) in this case as well.

Induction Step: We need to show that the lemma holds for \( i = k + 1 \). If both processes retrieve \( p.cseq[k+1].cfg \) and \( q.cseq[k+1].cfg \) through consensus, then both \( p \) and \( q \) run consensus over the previous configuration. Since according to our hypothesis \( c_{\sigma_1}^p[k] = c_{\sigma_2}^q[k] \) then both processes will receive the same decided value, say \( c_{k+1} \), and hence \( p.cseq[k+1].cfg = q.cseq[k+1].cfg = c_{k+1} \). Similar to the base case, a server in \( c_k.Servers \) only receives the configuration \( c_{k+1} \) decided by the consensus instance run over \( c_k \). So processes \( p \) and \( q \) can only receive \( c_{k+1} \) from some server in \( c_k.Servers \) so they can only assign \( p.cseq[k+1].cfg = q.cseq[k+1].cfg = c_{k+1} \) at Line Alg. 539.

That completes the proof.

Lemma. 48 In any execution \( \xi \) of the algorithm, If for any process \( p \in I \), \( c_{\sigma}^p[i] \neq \bot \) in some state \( \sigma \) in \( \xi \), then \( c_{\sigma'}^p[i] \neq \bot \) in any state \( \sigma' \) that appears after \( \sigma \) in \( \xi \).

Proof. A value is assigned to \( c_{\sigma}^p[i] \) either after the invocation of a consensus instance, or while executing the read-config action. Since any configuration proposed for installation cannot be \( \bot \) (A37), and since there is at least one configuration proposed in the consensus instance (the one from \( p \)), then by the validity of the consensus service the decision will be a configuration \( c \neq \bot \). Thus, in this case \( c_{\sigma}^p[i] \) cannot be \( \bot \). Also in the read-config procedure, \( c_{\sigma}^p[i] \) is assigned to a value different than \( \bot \) according to Line A519. Hence, if \( c_{\sigma}^p[i] \neq \bot \) at state \( \sigma \) then it cannot become \( \bot \) in any state \( \sigma' \) after \( \sigma \) in execution \( \xi \).

Lemma. 49 Let \( \sigma_1 \) be some state in an execution \( \xi \) of the algorithm. Then for any process \( p \), if \( k = \max\{i : c_{\sigma_1}^p[i] \neq \bot \} \), then \( c_{\sigma_1}^p[j] \neq \bot \), for \( 0 \leq j < k \).

Proof. Let us assume to derive contradiction that there exists \( j < k \) such that \( c_{\sigma_1}^p[j] = \bot \) and \( c_{\sigma_1}^p[j + 1] \neq \bot \). Suppose w.l.o.g. that \( j = k − 1 \) and that \( \sigma_1 \) is the state immediately after the
assignment of a value to \( c_{\sigma_1}^p[k] \), say \( c_k \). Since \( c_{\sigma_1}^p[k] \neq \bot \), then \( p \) assigned \( c_k \) to \( c_{\sigma_1}^p[k] \) in one of the following cases: (i) \( c_k \) was the result of the consensus instance, or (ii) \( p \) received \( c_k \) from a server during a read-config action. The first case is trivially impossible as according to Lemma 45 \( p \) decides for \( k \) when it runs consensus over configuration \( c_{\sigma_1}^p[k-1]. cfg \). Since this is equal to \( \bot \), then we cannot run consensus over a non-existent set of processes. In the second case, \( p \) assigns \( c_{\sigma_1}^p[k] = c_k \) in Line A49. The value \( c_k \) was however obtained when \( p \) invoked get-next-config on configuration \( c_{\sigma_1}^p[k-1].cfg \). In that action, \( p \) sends \( \text{READ-CONFIG} \) messages to the servers in \( c_{\sigma_1}^p[k-1].cfg \). Servers and waits until a quorum of servers replies. Since we assigned \( c_{\sigma_1}^p[k] = c_k \) it means that get-next-config terminated at some state \( \sigma' \) before \( \sigma_1 \) in \( c_k \), and thus: (a) a quorum of servers in \( c_{\sigma_1}^p[k-1].cfg \). Servers replied, and (b) there exists a server \( s \) among those that replied with \( c_k \). According to our assumption however, \( c_{\sigma_1}^p[k-1] = \bot \) at \( \sigma_1 \). So if state \( \sigma' \) is before \( \sigma_1 \) in \( c_k \), then by Lemma 48 it follows that \( c_{\sigma_1}^p[k-1] = \bot \). This however implies that \( p \) communicated with an empty configuration, and thus no server replied to \( p \). This however contradicts the assumption that a server replied with \( c_k \) to \( p \).

Since any process traverses the configuration sequence starting from the initial configuration \( c_0 \), then with a simple induction we can show that \( c_{\sigma_1}^p[j] \neq \bot \), for \( 0 \leq j \leq k \). □

We can now move to an important lemma that shows that any read-config action returns an extension of the configuration sequence returned by any previous read-config action. First, we show that the last finalized configuration observed by any read-config action is at least as recent as the finalized configuration observed by any subsequent read-config action.

**Lemma.** 50 If at a state \( \sigma \) of an execution \( \xi \) of the algorithm, if \( \mu(c_{\sigma}^p) = k \) for some process \( p \), then for any element \( 0 \leq j < k \), \( \exists Q \in c_{\sigma}^p[j].cfg \) Quorums such that \( \forall s \in Q, \sigma.s.nextC(c_{\sigma}^p[j].cfg) = c_{\sigma}^p[j+1] \).

**Proof.** This lemma follows directly from the algorithm. Notice that whenever a process assigns a value to an element of its local configuration (Lines A47 and A17), it then propagates this value to a quorum of the previous configuration (Lines A48 and A18). So if a process \( p \) assigned \( c_j \) to an element \( c_{\sigma}^p[j] \) in some state \( \sigma' \) in \( \xi \), then \( p \) may assign a value to the \( j+1 \) element of \( c_{\sigma'}^p[j+1] \) only after put-config(\( c_{\sigma'}^p[j].cfg, c_{\sigma}^p[j] \)) occurs. During put-config action, \( p \) propagates \( c_{\sigma}^p[j] \) in a quorum \( Q \in c_{\sigma'}^p[j+1].cfg \). Quorums. Hence, if \( c_{\sigma}^p[k] \neq \bot \), then \( p \) propagated each \( c_{\sigma}^p[j] \), for \( 0 \leq j \leq k \) to a quorum of servers \( Q \in c_{\sigma'}^p[j+1].cfg \). Quorums. And this completes the proof. □

**Lemma.** 51 (Configuration Prefix) Let \( \pi_1 \) and \( \pi_2 \) two completed read-config actions invoked by processes \( p_1, p_2 \in I \) respectively, such that \( \pi_1 \rightarrow \pi_2 \) in an execution \( \xi \). Let \( \sigma_1 \) be the state after the response step of \( \pi_1 \) and \( \sigma_2 \) the state after the response step of \( \pi_2 \). Then \( c_{\sigma_1}^{p_1} \leq_p c_{\sigma_2}^{p_2} \).

**Proof.** Let \( \nu_1 = \nu(c_{\sigma_1}^{p_1}) \) and \( \nu_2 = \nu(c_{\sigma_2}^{p_2}) \). By Lemma 47 for any \( i \) such that \( c_{\sigma_1}^{p_1}[i] \neq \bot \) and \( c_{\sigma_2}^{p_2}[i] \neq \bot \), then \( c_{\sigma_1}^{p_1}[i].cfg = c_{\sigma_2}^{p_2}[i].cfg \). Also from Lemma 49 we know that for \( 0 \leq j \leq \nu_2 \), \( c_{\sigma_2}^{p_2}[j] \neq \bot \). So if we can show that \( \nu_1 \leq \nu_2 \) then the lemma follows.

Let \( \mu = \mu(c_{\sigma_2}^{p_2}) \) be the last finalized element which \( p_2 \) established in the beginning of the read-config action \( \pi_2 \) (Line A52) at some state \( \sigma' \) before \( \sigma_2 \). It is easy to see that \( \mu \leq \nu_2 \). If \( \nu_1 \leq \mu \) then \( \nu_1 \leq \nu_2 \) and the lemma follows. Thus, it remains to examine the case where \( \mu < \nu_1 \). Notice that since \( \pi_1 \rightarrow \pi_2 \) then \( \sigma_1 \) appears before \( \sigma' \) in execution \( \xi \). By Lemma 50, we know that by \( \sigma_1 \), \( \exists Q \in c_{\sigma_1}^{p_1}[j].cfg \). Quorums, for \( 0 \leq j < \nu_1 \), such that \( \forall s \in Q, \sigma.s.nextC = c_{\sigma_1}^{p_1}[j+1] \). Since \( \mu < \nu_1 \), then it must be the case that \( \exists Q \in c_{\sigma_1}^{p_1}[\mu].cfg \). Quorums such that \( \forall s \in Q, \sigma.s.nextC = c_{\sigma_1}^{p_1}[\mu+1] \).
But by Lemma \[47\] we know that $c_{\sigma_1}^p[\mu].cfg = c_{\sigma_2}^p[\mu].cfg$. Let $Q'$ be the quorum that replies to the read-next-config occurred in $p_2$, on configuration $c_{\sigma_2}^p[\mu].cfg$. By definition $Q \cap Q' \neq \emptyset$, thus there is a server $s \in Q \cap Q'$ that sends $s.nextC = c_{\sigma_1}^p[\mu + 1]$ to $p_2$ during $\pi_2$. Since $c_{\sigma_1}^p[\mu + 1] \neq \bot$ then $p_2$ assigns $c_{\sigma_2}^p[\mu + 1] = c_{\sigma_1}^p[\mu + 1]$, and repeats the process in the configuration $c_{\sigma_2}^p[\mu + 1].cfg$. Since every configuration $c_{\sigma_2}^p[j].cfg$, for $\mu \leq j < \nu_1$, has a quorum of servers with $s.nextC$, then by a simple induction it can be shown that the process will be repeated for at least $\nu_1 - \mu$ iterations, and every configuration $c_{\sigma_2}^p[j] = c_{\sigma_1}^p[j]$, at some state $\sigma''$ before $\sigma_2$. Thus, $c_{\sigma_2}^p[j] = c_{\sigma_1}^p[j]$, for $0 \leq j \leq \nu_1$. Hence $\nu_1 \leq \nu_2$ and the lemma follows in this case as well.

**Lemma.** \[52\] Let $\sigma$ and $\sigma'$ two states in an execution $\xi$ such that $\sigma$ appears before $\sigma'$ in $\xi$. Then for any process $p$ must hold that $\mu(c_{\sigma_1}^p) \leq \mu(c_{\sigma_2}^p)$.

**Proof.** This lemma follows from the fact that if a configuration $k$ is such that $c_{\sigma_1}^p[k].status = F$ at a state $\sigma$, then $p$ will start any future read-config action from a configuration $c_{\sigma_2}^p[j].cfg$ such that $j \geq k$. But $c_{\sigma_2}^p[j].cfg$ is the last finalized configuration at $\sigma'$ and hence $\mu(c_{\sigma_2}^p) \geq \mu(c_{\sigma_1}^p)$.

**Lemma.** \[53\] (Configuration Progress) Let $\pi_1$ and $\pi_2$ two completed read-config actions invoked by processes $p_1, p_2 \in \mathcal{P}$ respectively, such that $\pi_1 \rightarrow \pi_2$ in an execution $\xi$. Let $\sigma_1$ be the state after the response step of $\pi_1$ and $\sigma_2$ the state after the response step of $\pi_2$. Then $\mu(c_{\sigma_1}^{p_1}) \leq \mu(c_{\sigma_2}^{p_2})$.

**Proof.** By Lemma \[51\] it follows that $c_{\sigma_1}^{p_1}$ is a prefix of $c_{\sigma_2}^{p_2}$. Thus, if $\nu_1 = \nu(c_{\sigma_1}^{p_1})$ and $\nu_2 = \nu(c_{\sigma_2}^{p_2})$, $\nu_1 \leq \nu_2$. Let $\mu_1 = \mu(c_{\sigma_1}^{p_1})$, such that $\mu_1 \leq \nu_1$, be the last element in $c_{\sigma_1}^{p_1}$, where $c_{\sigma_1}^{p_1}[\mu_1].status = F$. Let now $\mu_2 = \mu(c_{\sigma_2}^{p_2})$, be the last element which $p_2$ obtained in Line A4 during $\pi_2$ such that $c_{\sigma_2}^{p_2}[\mu_2].status = F$ in some state $\sigma'$ before $\sigma_2$. If $\mu_2 \geq \mu_1$, and since $\sigma_2$ is after $\sigma'$, then by Lemma \[52\] $\mu_2 \leq \mu(c_{\sigma_2}^{p_2})$ and hence $\mu_1 \leq \mu(c_{\sigma_2}^{p_2})$ as well.

It remains to examine the case where $\mu_2 < \mu_1$. Process $p_1$ sets the status of $c_{\sigma_1}^{p_1}[\mu_1]$ to $F$ in two cases: (i) either when finalizing a reconfiguration, or (ii) when receiving an $s.nextC = \langle c_{\sigma_1}^{p_1}[\mu_1].cfg, F \rangle$ from some server $s$ during a read-config action. In both cases $p_1$ propagates the $\langle c_{\sigma_1}^{p_1}[\mu_1].cfg, F \rangle$ to a quorum of servers in $c_{\sigma_1}^{p_1}[\mu_1 - 1].cfg$ before completing. We know by Lemma \[51\] that since $\pi_1 \rightarrow \pi_2$ then $c_{\sigma_1}^{p_1}$ is a prefix in terms of configurations of the $c_{\sigma_2}^{p_2}$. So it must be the case that $\mu_2 < \mu_1 \leq \nu(c_{\sigma_2}^{p_2})$. Thus, during $\pi_2$, $p_2$ starts from the configuration at index $\mu_2$ and in some iteration performs get-next-config in configuration $c_{\sigma_2}^{p_2}[\mu_1 - 1]$. According to Lemma \[47\] $c_{\sigma_2}^{p_2}[\mu_1 - 1].cfg = c_{\sigma_2}^{p_2}[\mu_1 - 1].cfg$. Since $\pi_1$ completed before $\pi_2$, then it must be the case that $\sigma_1$ appears before $\sigma'$ in $\xi$. However, $p_2$ invokes the get-next-config operation in a state $\sigma''$ which is either equal to $\sigma'$ or appears after $\sigma'$ in $\xi$. Thus, $\sigma''$ must appear after $\sigma_1$ in $\xi$. From that it follows that when the get-next-config is executed by $p_2$ there is already a quorum of servers in $c_{\sigma_2}^{p_2}[\mu_1 - 1].cfg$, say $Q_1$, that received $\langle c_{\sigma_1}^{p_1}[\mu_1].cfg, F \rangle$ from $p_1$. Since, $p_2$ waits from replies from a quorum of servers from the same configuration, say $Q_2$, and since the nextConfig variable at each server is monotonic (Lemma \[46\]), then there is a server $s \in Q_1 \cap Q_2$, such that $s$ replies to $p_2$ with $s.nextC = \langle c_{\sigma_1}^{p_1}[\mu_1].cfg, F \rangle$. So, $c_{\sigma_2}^{p_2}[\mu_1]$ gets $\langle c_{\sigma_2}^{p_2}[\mu_1].cfg, F \rangle$, and hence $\mu(c_{\sigma_2}^{p_2}) \geq \mu_1$ in this case as well. This completes our proof.

**Lemma.** \[17\] Let $\pi$ be a complete reconfiguration operation by a reconfigurer $rc$ in an execution $\xi$ of ARES. If $\sigma_1$ is the state in $\xi$ following the termination of the read-config action during $\pi$, then $\pi$ invokes a finalize-config($c_{\sigma_2}^{rc}$) at a state $\sigma_2$ in $\xi$, with $\nu(c_{\sigma_2}^{rc}) = \nu(c_{\sigma_1}^{rc}) + 1$. 

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Proof. This lemma follows directly from the implementation of the reconfig operation. Let \( \rho \) be a reconfiguration operation \( \text{reconfig}(c) \). At first, \( \rho \) invokes a read-config to retrieve a latest value of the global configuration sequence, \( c^r_{\sigma_1} \), in a state \( \sigma_1 \) in \( \xi \). During the add-config action, \( \rho \) proposes the addition of \( c \), and appends at the end of \( c^r_{\sigma_1} \) the decision \( d \) of the consensus protocol. Therefore, if \( c^r_{\sigma_1} \) is extended by \( \langle d, P \rangle \) (Line Alg. 51), and hence the add-config action returns a configuration sequence \( c^r_{\sigma_1'} \) with length \( \nu(c^r_{\sigma_1'}) = \nu(c^r_{\sigma_1}) + 1 \). As \( \nu(c^r_{\sigma_1'}) \) does not change during the update-config action, then \( c^r_{\sigma_1'} \) is passed to the finalize-config action at state \( \sigma_2 \), and hence \( c^r_{\sigma_2} = c^r_{\sigma_1} \). Thus, \( \nu(c^r_{\sigma_2}) = \nu(c^r_{\sigma_1'}) = \nu(c^r_{\sigma_1}) + 1 \) and the lemma follows.

Lemma. \[ \text{Suppose } \xi \text{ is an execution of ARES. For any state } \sigma \text{ in } \xi, \text{ if } c^p_\sigma[j], \text{ status } = F \text{ for some process } p \in \mathcal{I}, \text{ then there exists a reconfig operation } \rho \text{ by a reconfigurer } rc \in G, \text{ such that (i) } rc \text{ invokes finalize-config}(c^c_{\sigma_1'}) \text{ during } \rho \text{ at some state } \sigma' \text{ in } \xi, (ii) } \nu(c^c_{\sigma_1'}) = j, \text{ and (iii) } \mu(c^c_{\sigma_1'}) < j. \]

Proof. A process sets the status of a configuration \( c \) to \( F \) in two cases: (i) either during a finalize-config(seq) action such that \( \nu(\text{seq}) = \langle c, P \rangle \) (Line Alg. 33), or (ii) when it receives \( \langle c, F \rangle \) from a server \( s \) during a read-next-config action. Server \( s \) sets the status of a configuration \( c \) to \( F \) only if it receives a message that contains \( \langle c, F \rangle \) (Line Alg. 10). So, (ii) is possible only if \( c \) is finalized during a reconfig operation.

Let, w.l.o.g., \( \rho \) be the first reconfiguration operation that finalizes \( c^p_\sigma[j].cfg \). To do so, process \( rc \) invokes finalize-config(c^c_{\sigma_1'}) during \( \rho \), at some state \( \sigma' \) that appears before \( \sigma \) in \( \xi \). By Lemma 47 \( c^p_\sigma[j].cfg = c^c_{\sigma_1'}[j].cfg \). Since, \( rc \) finalizes \( c^c_{\sigma_1'}[j] \), then this is the last entry of \( c^c_{\sigma_1'} \) and hence \( \nu(c^c_{\sigma_1'}) = j \). Also, by Lemma 18 it follows that the read-config action of \( \rho \) returned a configuration \( c^c_{\sigma_1'} \) in some state \( \sigma'' \) that appeared before \( \sigma' \) in \( \xi \), such that \( \nu(c^c_{\sigma_1''}) < \nu(c^c_{\sigma_1'}) \). Since by definition, \( \mu(c^c_{\sigma_1}) \leq \nu(c^c_{\sigma_1'}) \), then \( \mu(c^c_{\sigma_1''}) < j \). However, since only \( \langle c, P \rangle \) is added to \( c^c_{\sigma_1''} \) to result in \( c^c_{\sigma_1} \), then \( \mu(c^c_{\sigma_1}) = \mu(c^c_{\sigma_1'}) \). Therefore, \( \mu(c^c_{\sigma_1}) < j \) as well and the lemma follows.

Lemma. \[ \text{Let } \pi_1 \text{ and } \pi_2 \text{ be two completed read/write/reconfig operations invoked by processes } p_1 \text{ and } p_2 \text{ in } \mathcal{I}, \text{ in an execution } \xi \text{ of ARES, such that, } \pi_1 \rightarrow \pi_2. \text{ If } c^1._1.put-data(\langle \tau_{\pi_1}, v_{\pi_1} \rangle) \text{ is the last put-data action of } \pi_1 \text{ and } \sigma_2 \text{ is the state in } \xi, \text{ after the completion of the first read-config action of } \pi_2, \text{ then there exists a } c^2._2.put-data(\langle \tau_{\pi_2}, v \rangle) \text{ action in some configuration } c_2 = c^2_{\sigma_2}[i].cfg, \text{ for } \mu(c^2_{\sigma_2}) \leq i \leq \nu(c^2_{\sigma_2}), \text{ such that (i) it completes in a state } \sigma' \text{ before } \sigma_2 \text{ in } \xi, \text{ and (ii) } \tau \geq \tau_{\pi_1}. \]

Proof. Note that from the definitions of \( \nu(\cdot) \) and \( \mu(\cdot) \), we have \( \mu(c^p_{\sigma_2}) \leq \nu(c^p_{\sigma_2}). \) Let \( \sigma_1 \) be the state in \( \xi \) after the completion of \( c^1._1.put-data(\langle \tau_{\pi_1}, v_{\pi_1} \rangle) \) and \( \sigma_1' \) be the state in \( \xi \) following the response step of \( \pi_1 \). Since any operation executes put-data on the last discovered configuration then \( c_1 \) is the last configuration found in \( c^p_{\sigma_1} \), and hence \( c_1 = \nu(c^p_{\sigma_1}) \). By Lemma 52 we have \( \mu(c^p_{\sigma_1}) \leq \mu(c^p_{\sigma_1'}) \) and by Lemma 53 we have \( \mu(c^p_{\sigma_1'}) \leq \mu(c^p_{\sigma_2}) \), since \( \pi_2 \) (and thus its first read-config action) is invoked after \( \sigma_1' \) (and thus after the last read-config action during \( \pi_1 \)). Hence, combining the two implies that \( \mu(c^p_{\sigma_1}) \leq \mu(c^p_{\sigma_2}). \) Now from the last implication and the first statement we have \( \mu(c^p_{\sigma_1}) \leq \nu(c^p_{\sigma_2}). \) Therefore, we consider two cases: (a) \( \mu(c^p_{\sigma_2}) \leq \nu(c^p_{\sigma_1}) \) and (b) \( \mu(c^p_{\sigma_2}) > \nu(c^p_{\sigma_1}) \). Let for abbreviation denote by \( \mu_i = \mu(c^p_{\sigma_i}) \), and \( \nu_i = \nu(c^p_{\sigma_i}) \).

Case (a): Since \( \pi_1 \rightarrow \pi_2 \) then, by Lemma 51 \( c^p_{\sigma_2} \) value returned by read-config at \( p_2 \) during the execution of \( \pi_2 \) satisfies \( c^p_{\sigma_1} \leq_p c^p_{\sigma_2} \). Therefore, \( \nu(c^p_{\sigma_1}) \leq \nu(c^p_{\sigma_2}) \), and in this case \( \mu(c^p_{\sigma_2}) \leq \nu(c^p_{\sigma_1}) \). Since \( c_1 \) is the last configuration in \( c^p_{\sigma_1} \), then it has index \( \nu(c^p_{\sigma_1}) \). So if we take
\( c_2 = c_1 \) then the \( c_1\text{-put-data}(\langle \tau_{\pi_1}, v_{\pi_1} \rangle) \) action trivially satisfies both conditions as: (i) it completes in state \( \sigma_1 \) which appears before \( \sigma_2 \), and (ii) it puts a pair \( \langle \tau, v \rangle \) such that \( \tau = \tau_{\pi_1} \).

**Case (b):** Note that there exists a reconfiguration client \( rc \) that invokes reconfig operation \( \rho \), during which it executes the finalize-config(\( c^e \)) that finalized configuration with index \( \nu(c^e) = \mu(c^p_2) \).

Let \( \sigma \) be the state immediately after the read-config of \( \rho \). Now, we consider two sub-cases: (i) \( \sigma \) appears before \( \sigma_1 \) in \( \xi \), or (ii) \( \sigma \) appears after \( \sigma_1 \) in \( \xi \).

Subcase (b)(i): Since read-config at \( \sigma \) completes before the invocation of last read-config of operation \( \pi_1 \) then, either \( c^e_\sigma \prec_p c^p_1 \) or \( c^e_\sigma = c^p_1 \) due to Lemma \[51\]. Suppose \( c^e_\sigma \prec_p c^p_1 \), then according to Lemma \[17\] \( rc \) executes finalize-config on configuration sequence \( c^e \) with \( \nu(c^e) = \nu(c^e_\sigma) + 1 \). Since \( \nu(c^e_\sigma) = \mu(c^p_1) \), then \( \mu(c^p_1) = \nu(c^e_\sigma) + 1 \). If however, \( c^e_\sigma \prec_p c^p_1 \), then \( \nu(c^e_\sigma) < \nu(c^p_1) \) and thus \( \nu(c^e_\sigma) + 1 \leq \nu(c^p_1) \). This implies that \( \mu(c^p_1) \leq \nu(c^p_1) \) which contradicts our initial assumption for this case that \( \mu(c^p_2) > \nu(c^p_1) \). So this sub-case is impossible.

Now suppose, that \( c^e_\sigma = c^p_1 \). Then it follows that \( \nu(c^e_\sigma) = \nu(c^p_1) \), and that \( \mu(c^p_1) = \nu(c^p_1) + 1 \) in this case. Since \( \sigma_1 \) is the state after the last put-data during \( \pi_1 \), then if \( \sigma_2 \) is the state after the completion of the last read-config of \( \pi_1 \) (which follows the put-data), it must be the case that \( c^p_1 = c^p_1 \). So, during its last read-config process \( p_1 \) does not read the configuration indexed at \( \nu(c^p_1) + 1 \). This means that the put-config completes in \( \rho \) at state \( \sigma_\rho \) after \( \sigma_2 \) and the update-config operation is invoked at state \( \sigma_2 \) after \( \sigma_\rho \) with a configuration sequence \( c^e_\rho \). During the update operation \( \rho \) invokes get-data operation in every configuration \( c^e_\rho[i].cfg \) for \( \mu(c^e_\rho) \leq i \leq \nu(c^e_\rho) \). Notice that \( \nu(c^e_\rho) = \mu(c^p_2) = \nu(c^p_1) + 1 \) and moreover the last configuration of \( c^e_\rho \) was just added by \( \rho \) and it is not finalized. From this it follows that \( \mu(c^e_\rho) < \nu(c^e_\rho) \), and hence \( \mu(c^e_\rho) \leq \nu(c^p_1) \).

Therefore, \( \rho \) executes get-data in configuration \( c^e_\rho[j].cfg \) for \( j = \nu(c^p_1) \). Since \( p_1 \) invoked put-data(\( \langle \tau_{\pi_1}, v_{\pi_1} \rangle \)) at the same configuration \( \nu(c^p_1) \), and completed in a state \( \sigma_1 \) before \( \sigma_2 \), then by property \( \text{C1} \) of Definition \[31\] it follows that the get-data action will return a tag \( \tau \geq \tau_{\pi_1} \). Therefore, the maximum tag that \( \rho \) discovers is \( \tau_{\text{max}} \geq \tau \geq \tau_{\pi_1} \). Before invoking the finalize-config action, \( \rho \) invokes \( \nu(c^e_\rho).\text{put-data}(\langle \tau_{\text{max}}, v_{\text{max}} \rangle) \). Since \( \nu(c^e_\rho) = \mu(c^p_2) \), and since by Lemma \[47\] then the action put-data is invoked in a configuration \( c_2 = c^p_2[j].cfg \) such that \( j = \nu(c^p_1) \). Since the read-config action of \( \pi_2 \) observed configuration \( \mu(c^p_2) \), then it must be the case that \( \sigma_2 \) appears after the state in which the finalize-config was invoked and therefore after the state of the completion of the put-data action during \( \rho \). Thus, in this case both properties are satisfied and the lemma follows.

Subcase (b)(ii): Suppose in this case that \( \sigma \) occurs in \( \xi \) after \( \sigma_1 \). In this case the last put-data in \( \pi_1 \) completes before the invocation of the read-config in \( \rho \) in execution \( \xi \). Now we can argue recursively, \( \rho \) taking the place of operation \( \pi_2 \), that \( \mu(c^e_\rho) \leq \nu(c^p_1) \) and therefore, we consider two cases: (a) \( \mu(c^e_\rho) \leq \nu(c^p_1) \) and (b) \( \mu(c^e_\rho) > \nu(c^p_1) \). Note that there are finite number of operations invoked in \( \xi \) before \( \pi_2 \) is invoked, and hence the statement of the lemma can be shown to hold by a sequence of inequalities.

\[ \square \]

**Lemma.** \[20\] Let \( \pi_1 \) and \( \pi_2 \) denote completed read/write operations in an execution \( \xi \), from processes \( p_1, p_2 \in \mathcal{P} \) respectively, such that \( \pi_1 \rightarrow \pi_2 \). If \( \tau_{\pi_1} \) and \( \tau_{\pi_2} \) are the local tags at \( p_1 \) and \( p_2 \) after the completion of \( \pi_1 \) and \( \pi_2 \) respectively, then \( \tau_{\pi_1} \leq \tau_{\pi_2} \), if \( \pi_1 \) is a write operation then \( \tau_{\pi_1} < \tau_{\pi_2} \).

**Proof.** Let \( \langle \tau_{\pi_1}, v_{\pi_1} \rangle \) be the pair passed to the last put-data action of \( \pi_1 \). Also, let \( \sigma_2 \) be the
state in $\xi$ that follows the completion of the first read-config action during $\pi_2$. Notice that $\pi_2$ executes a loop after the first read-config operation and performs $c$.get-data (if $\pi_2$ is a read) or $c$.get-tag (if $\pi_2$ is a write) from all $c = c'_{p_2}$.cfg, for $\mu(c'_{p_2}) \leq i \leq \nu(c'_{p_2})$. By Lemma 19 there exists a $c'$.put-data($\langle \tau, v \rangle$) action by some operation $\pi'$ on some configuration $c' = c_{p_2}^2$.cfg, for $\mu(c_{p_2}^2) \leq j \leq \nu(c_{p_2}^2)$, that completes in some state $\sigma'$ that appears before $\sigma_2$ in $\xi$. Thus, the get-data or get-tag invoked by $p_2$ on $c_{p_2}^2$.cfg, occurs after state $\sigma_2$ and thus after $\sigma'$. Since the DAP primitives used satisfy properties C1 and C2 of Definition 31, then the get-tag action will return a tag $\tau_{p_2}'$ or a get-data action will return a pair $\langle \tau_{p_2}', v_{p_2}' \rangle$, with $\tau_{p_2}' \geq \tau$. As $p_2$ gets the maximum of all the tags returned, then by the end of the loop $p_2$ will retrieve a tag $\tau_{max} \geq \tau_{p_2}' \geq \tau \geq \tau_{p_1}$.

If now $\pi_2$ is a read, it returns $\langle \tau_{max}, v_{max} \rangle$ after propagating that value to the last discovered configuration. Thus, $\tau_{p_2} \geq \tau_{p_1}$. If however $\pi_2$ is a write, then before propagating the new value the writer increments the maximum timestamp discovered (Line A7,13) generating a tag $\tau_{p_2} \geq \tau_{max}$.

Therefore the operation $\pi_2$ propagates a tag $\tau_{p_2} > \tau_{p_1}$ in this case.

\section*{D Latency Analysis}

In this section we examine closely the latencies of each operation in ARES, and question the termination of each operation under various environmental conditions. For our analysis we assume that each operation inflicts constant computation overhead and delays are only introduced due to the message exchange among the processes. We will measure delays in time units of some global clock $T$. No process has access to $T$ and the clock can only be accessed by an external viewer. Let $d$ denote the minimum message delivery delay between any two processes in the service; let $D$ be the maximum delivery delay. Also, let $T(\pi)$ denote the communication delay of an operation (or action) $\pi$.

Given the message delivery bounds $d$ and $D$, and through inspection of the algorithm we can provide the delay bounds of the operations and actions used by ARES as follows:

\textbf{Lemma 55.} If any message send from a process $p_1$ to a process $p_2$, s.t. $p_1, p_2 \in I \cup S$, takes at least $d$ and at most $D$ time units to be delivered, then the following operations may terminate within the following time intervals: (i) $2d \leq T(\text{put-config}) \leq 2D$ and (ii) $2d \leq T(\text{read-next-config}) \leq 2D$.

From Lemma 55 we can derive the delay of a read-config action.

\textbf{Lemma 56.} For any read-config action $\phi$ such that it accepts an input $seq$ and returns $seq'$, if $\mu = \mu(\text{seq})$ and $\nu = \nu(\text{seq'})$ then $\phi$ takes:

$$4d(\nu - \mu + 1) \leq T(\phi) \leq 4D(\nu - \mu + 1)$$

(1)

From Lemma 56 it is apparent that the latency of a read-config action highly depends on the number of configurations installed, following the last finalized configuration as that was known by the process at the invocation of the read-config action. Let $\lambda = \nu - \mu$ denote the number of newly installed configurations. Now let us examine when a new configuration gets inserted in the configuration sequence by a reconfig operation. By ARES a reconfig operation has four phases: (i) reads the latest configuration sequence, (ii) adds the new configuration at the end of the sequence, (iii) transfers the knowledge to the added configuration, and (iv) finalizes the added configuration.
So, a new configuration is appended to the end of the configuration sequence (and it becomes visible to any operation) during the add-config action. In turn, the add-config action, runs a consensus algorithm to decide on the added configuration and then invokes a put-config action to add the decided configuration. Any operation that is invoked after the put-config action will observed the newly added configuration.

Notice that when multiple reconfigurations are invoked concurrently, then it might be the case that all participate to the same consensus instance and the configuration sequence is appended by a single configuration. The worst case scenario happens when all concurrent reconfigurations manage to append the configuration sequence by their configuration. In brief, this is possible when the read-config action of each reconfig operation appears after the put-config action of another reconfig operation.

![Figure 2: Successful reconfig operations.](image)

More formally we can build an execution where all reconfig operations append their configuration in the configuration sequence. Consider a partial execution $\xi$ that ends in a state $\sigma$. Suppose that every process $p \in I$ knows the same configuration sequence, $c_\sigma^p = c_\sigma$. Also let the last finalized operation in $c_\sigma$ be the last configuration of the sequence, e.g. $\mu(c_\sigma) = \nu(c_\sigma)$. Notice that $c_\sigma$ can also be the initial configuration sequence $c_\sigma^p_{\sigma_0}$. We extend $\xi_0$ by a series of reconfig operations, such that each reconfiguration $rc_i$ is invoked by a reconfigurer $r_i$ and attempts to add a configuration $c_i$. Let $rc_1$ be the first reconfiguration that performs the following actions without being concurrent with any other reconfig operation:

- read-config starting from $\mu(c_\sigma)$
- add-config completing both the consensus proposing $c_1$ and the put-config action writing the decided configuration

Since $rc_1$ its not concurrent with any other reconfig operation, then is the only process to propose a configuration in $\mu(c_\sigma)$, and hence by the consensus algorithm properties, $c_1$ is decided. Thus, $c_\sigma$ is appended by a tuple $(c_1, P)$.

Let now reconfiguration $rc_2$ be invoked immediately after the completion of the add-config action from $rc_1$. Since the local sequence at the beginning of $rc_2$ is equal to $c_\sigma$, then the read-config
action of $rc_2$ will also start from $\mu(c_\sigma)$. Since, $rc_1$ already propagated $c_1$ to $\mu(c_\sigma)$ during is put-config action, then $rc_2$ will discover $c_1$ during the first iteration of its read-config action, and thus it will repeat the iteration on $c_1$. Configuration $c_1$ is the last in the sequence and thus the read-config action of $rc_2$ will terminate after the second iteration. Following the read-config action, $rc_2$ attempts to add $c_2$ in the sequence. Since $rc_1$ is the only reconfiguration that might be concurrent with $rc_2$, and since $rc_1$ already completed consensus in $\mu(c_\sigma)$, then $rc_2$ is the only operation to run consensus in $c_1$. Therefore, $c_2$ is accepted and $rc_2$ propagates $c_2$ in $c_1$ using a put-config action.

So in general we let configuration $rc_i$ to be invoked after the completion of the add-config action from $rc_{i-1}$. As a result, the read-config action of $rc_i$ performs $i$ iterations, and the configuration $c_i$ is added immediately after configuration $c_{i-1}$ in the sequence. Figure 2 illustrates our execution construction for the reconfiguration operations.

It is easy to notice that such execution results in the worst case latency for all the th reconfiguration operations $rc_1, rc_2, \ldots, rc_i$. We can now compute the minimum latency we need to add $k$ new configurations in the configuration sequence starting from the state $\sigma$ of execution $\xi$. For simplicity of our analysis we assume that any consensus instance takes the same time to terminate and that is $T(CN)$.

**Lemma 57.** Starting from the last state of $\xi$, $\sigma$, and given that $d$ is the minimum communication delay, then $k$ configurations can be appended to $c_\sigma$, in time: $T(k) \geq 4d \sum_{i=1}^{k} i + k (T(CN) + 2d)$ in our execution construction.

*Proof.* Figure 2 shows the timings of each reconfiguration operation. In particular, consider the first reconfiguration $rc_1$. During its read-config $rc_1$ does not discover new configurations and thus, if $seq_1$ is the input and $seq'_1$ the output configuration, $\mu(seq_1) = \nu(seq'_1)$. Thus, by Lemma 56, the read-config takes at least time $4d$. Since the consensus algorithm takes $T(CN)$ and the put-config action at least $2d$ (see Lemma 55), then $rc_1$ takes time at least:

$$T(1) \geq 4d + T(CN) + 2d$$

to install configuration $c_1$. Reconfiguration $rc_2$, executes two iterations during its read-config action, and $\nu(seq'_2) = \mu(seq_2) + 1$. Thus, by Lemma 56, the read-config of $rc_2$ takes at least time $4d * 2$. Until $rc_2$ installs $c_2$ it needs also time $T(CN)$ for the consensus algorithm and $2d$ for the put-config action. Hence, both configurations $c_1$ and $c_2$ are appended in time at least:

$$T(2) \geq T(1) + 8d + T(CN) + 2d$$

$$\geq 12d + 2(T(CN) + 2d)$$

So in general to install the configuration $k$ it takes time at least $4d * k + T(CN) + 2d$, and thus to append the sequence with all the configurations $c_1, \ldots, c_k$ it takes time at least:

$$T(k) \geq T(k-1) + 4kd + T(CN) + 2d \geq 4d \sum_{i=1}^{k} i + k(T(CN) + 2d)$$

And that completes our proof. \qed

Finally, we can compute the maximum time that a read/write operation needs before completing. From close examination of the algorithm, the DAPs used by ARES have an impact on the delay of a read and write operation. For our analysis we assume that all get-data, get-tag and put-data
primitives use two phases of communication as this capture the DAP we present in this work as well as the most common implementations of atomic registers.

**Lemma 58.** If any message send from a process \( p_1 \) to a process \( p_2 \), s.t. \( p_1, p_2 \in I \cup S \), takes at least \( d \) and at most \( D \) time units to be delivered, then the DAPs may terminate in: (i) \( 2d \leq T(\text{put-data}) \leq 2D \); (ii) \( 2d \leq T(\text{get-tag}) \leq 2D \); and (iii) \( 2d \leq T(\text{get-data}) \leq 2D \).

Having the delays for the DAPs we can now compute the delay of a read/write operation \( \pi \).

**Lemma 59.** Let \( \sigma_s \) and \( \sigma_e \) be the states before the invocation and after the completion step of a read/write operation \( \pi \) by \( p \) respectively, in some execution \( \xi \) of ARES,. Then \( \pi \) takes time at most:

\[
T(\pi) \leq 6D \left[ \nu(c_{\sigma_e}^p) - \mu(c_{\sigma_s}^p) + 2 \right]
\]

**Proof.** By algorithm examination we can see that any read/write operation performs the following actions in this order: (i) read-config, (ii) get-data/(or get-tag), (iii) put-data, and (iv) read-config. Let \( \sigma_1 \) be the state when the first read-config, denoted by read-config\(_1\), action terminates. By Lemma \ref{lemma:read-config} the action will take time:

\[
T(\text{read-config}_1) \leq 4D(\nu(c_{\sigma_1}^p) - \mu(c_{\sigma_s}^p) + 1)
\]

The get-data action that follows the read-config (Lines A7:34 A7:35) is also executed at most \( (\nu(c_{\sigma_1}^p) - \mu(c_{\sigma_s}^p) + 1) \) given that no new finalized configuration was discovered by the read-config action. Finally, the put-data and the second read-config actions of \( \pi \) may be invoked at most \( (\nu(c_{\sigma_e}^p) - \nu(c_{\sigma_1}^p) + 1) \) times, given that the read-config action discovers one new configuration every time it runs. Merging all the outcomes, the total time of \( \pi \) can be at most:

\[
T(\pi) \leq 4D(\nu(c_{\sigma_1}^p) - \mu(c_{\sigma_s}^p) + 1) + 2D(\nu(c_{\sigma_1}^p) - \mu(c_{\sigma_s}^p) + 1) + (4D + 2D)(\nu(c_{\sigma_e}^p) - \nu(c_{\sigma_1}^p) + 1)
\]

\[
\leq 6D \left[ \nu(c_{\sigma_e}^p) - \mu(c_{\sigma_s}^p) + 2 \right]
\]

And the lemma follows.

It remains now to examine if a read/write operation may “catch up” with any ongoing reconfigurations. For the sake of a worst case analysis we will assume that reconfiguration operations may communicate respecting the minimum delay \( d \), whereas read and write operations suffer the maximum delay \( D \) in each message exchange. We will split our analysis into three directions, with respect to the number of configurations installed \( k \), and the bound on the minimum delay \( d \): (i) \( k \) is finite, and \( d \) may be unbounded small; (ii) \( k \) is infinite, and \( d \) may be unbounded small; (iii) \( k \) is infinite, and \( d \) can be bounded.

**\( k \) is finite, and \( d \) may be unbounded small.** In this case we assume a finite number of installed configurations. Also, as the \( d \) is unbounded, it follows that reconfigurations may be installed almostinstantaneously. Let us first examine what is the maximum delay bound of a any read/write operation.

**\( k \) is infinite, and \( d \) is bounded.** We will compute the bound on \( d \) with respect to the \( D \) and the number of configurations to be installed \( k \) if we want to allow a read/write operation to reach ongoing reconfigurations.
Lemma 60. A read/write operation $\pi$ may terminate in any execution $\xi$ of ARES given that $k$ configurations are installed during $\pi$, if $d \geq \frac{3D}{k} - \frac{T(CN)}{2(k+2)}$.

Proof. By Lemma 57, $k$ configurations may be installed in at least: $T(k) \geq 4d \sum_{i=1}^{k} i + k(T(CN) + 2d)$. Also by Lemma 59, we know that operation $\pi$ takes at most $T(\pi) \leq 6D (\nu(c_{p_e}) - \mu(c_{p_s}) + 2)$. Assuming that $k = \nu(c_{p_e}) - \mu(c_{p_s})$, the total number of configurations observed during $\pi$, then $\pi$ may terminate before a $k + 1$ configuration is added in the configuration sequence if $6D(k + 2) \leq 4d \sum_{i=1}^{k} i + k(T(CN) + 2d)$ then we have $d \geq \frac{3D}{k} - \frac{T(CN)}{2(k+2)}$. And that completes the lemma.

5 Efficient state transfer during reconfiguration

Figure 3: An illustration of the recon client $rc$ transferring data from servers of configuration $C$ to those in $C'$. 