Asynchronous Reconfiguration with Byzantine Failures

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ABSTRACT
Replicated services are inherently vulnerable to failures and security breaches. In a long-running system, it is, therefore, indispensable to maintain a reconfiguration mechanism that would replace faulty replicas with correct ones. An important challenge is to enable reconfiguration without affecting the availability and consistency of the replicated data: the clients should be able to get correct service even when the set of service replicas is being updated.

In this paper, we address the problem of reconfiguration in the presence of Byzantine failures: faulty replicas or clients may arbitrarily deviate from their expected behavior. We describe a generic technique for building asynchronous and Byzantine fault-tolerant reconfigurable objects: clients can manipulate the object data and issue reconfiguration calls without reaching consensus on the current configuration. With the help of forward-secure digital signatures, our solution makes sure that superseded and possibly compromised configurations are harmless, that slow clients cannot be fooled into reading stale data, and that Byzantine clients cannot cause a denial of service by flooding the system with reconfiguration requests. Our approach is modular and based on dynamic Byzantine lattice agreement abstraction, and we discuss how to extend it to enable Byzantine fault-tolerant implementations of a large class of reconfigurable replicated services.

KEYWORDS
Reconfiguration, Asynchronous system, Byzantine faults

1 INTRODUCTION

Replication and quorums. Replication is a natural way to ensure availability of shared data in the presence of failures. A collection of replicas, each holding a version of the data, ensure that the clients get a desired service, even when some replicas become unavailable or hacked by a malicious adversary. Consistency of the provided service requires the replicas to synchronize: intuitively, every client should be able to operate on the most “up-to-date” data, regardless of the set of replicas it can reach.

It always makes sense to assume as little as possible about the environment in which a system we design is expected to run. For example, asynchronous distributed systems do not rely on timing assumptions, which makes them extremely robust with respect to communication disruptions and computational delays. It is, however, notoriously difficult and sometimes even impossible to make such systems fault-tolerant. The folklore CAP theorem [12, 22] states that no replicated service can combine consistency, availability, and partition-tolerance. In particular, no consistent and available read-write storage can be implemented in the presence of partitions: clients in one partition are unable to keep track of the updates taking place in another one.

Therefore, fault-tolerant storage systems tend to assume that partitions are excluded, e.g., by requiring a majority of replicas to be correct [6]. More generally, one can assume a quorum system, e.g., a set of subsets of replicas satisfying the intersection and availability properties [21]. Every (read or write) request from a client should be accepted by a quorum of replicas. As every two quorums have at least one replica in common, intuitively, no client can miss previously written data.

Of course, failures of replicas may jeopardize the underlying quorum system. In particular, we may find ourselves in a system in which no quorum is available and, thus, no operation may be able to terminate. Even worse, if the replicas are subject to Byzantine failures, we may not be able to guarantee the very correctness of read values.

Asynchronous reconfiguration. To anticipate such scenarios in a long run, we must maintain a reconfiguration mechanism that enables replacing compromised replicas with correct ones and update the corresponding quorum assumptions. A challenge here is to find an asynchronous implementation of reconfiguration in a system where both clients and replicas are subject to Byzantine failures that can be manifested by arbitrary and even malicious behavior. In the world of selfishly driven blockchain users, a reconfiguration mechanism must be prepared for this.

Recently, a number of reconfigurable systems were proposed for asynchronous crash-fault environments [2, 3, 20, 24, 27, 35] that were first applied to (read-write) storage systems [2, 3, 20], and then extended to max-registers [24, 35] and more general lattice data type [27].

These proposals tend to ensure that the clients reach a form of “loose” agreement on the currently active configurations, which can be naturally expressed via the lattice agreement abstraction [8, 17]. We allow clients to (temporarily) live in different worlds, as long as these worlds are properly ordered. For example, we may represent a configuration as a set of updates (additions and removals of replicas) and require that all installed configurations should be related by containment. A configuration becomes stale as soon as a new configuration representing a proper superset of updates is installed.

Challenges of Byzantine fault-tolerant reconfiguration. In this paper, we focus on Byzantine fault-tolerant reconfiguration. We have to address here several challenges, specific to dynamic systems with Byzantine faults, which does not allow to simply employ the existing crash fault-tolerant solutions.

First, when we build a system out of lower-level components, we need to make sure that the outputs provided by these components are “authentic”. Whenever a (potentially Byzantine) process claims to have obtained a value v (e.g., a new configuration estimate) from an underlying object (e.g., Lattice Agreement), it should
also provide a proof \( \sigma \) that can be independently verified by every correct process. The proof typically consists of digital signatures provided by a quorum of replicas of some configuration. We abstract this requirement out by equipping the object with a function \( \text{VerifyOutputValue} \) that returns a boolean value, provided \( v \) and \( \sigma \). When invoked by a correct process, the function returns \textit{true} if and only if \( v \) has indeed been produced by the object. When "chaining" the objects, i.e., adopting the output \( v \) provided by an object \( A \) as an input for another object \( B \), which is the typical scenario in our system, a correct process invokes \( A.\text{VerifyOutputValue}(v, \sigma) \), where \( \sigma \) is the proof associated with \( v \) by the implementation of \( A \). This way, only values actually produced by \( A \) can be used as inputs to \( B \).

Second, we face the "I still work here" attack [1]. It is possible that a client that did not log into the system for a long time tries to access a stale configuration in which some quorum is entirely compromised by the Byzantine adversary. The client can therefore be provided with an inconsistent view on the shared data. Thus, before accepting a new configuration, we need to make sure that the stale ones are no longer capable of processing data requests from the clients. We address this issue via the use of a forward-secure signature scheme [10]. Intuitively, every replica is provided with a distinct private key associated to each configuration. Before a configuration is replaced with a newer one, at least a quorum of its replicas are asked to destroy their private keys. Therefore, even if the replicas are to become Byzantine in the future, they will not be able to provide slow clients with inconsistent values. The stale configuration simply becomes non-responsive, as in crash-fault-tolerant reconfigurable systems.

Unfortunately, in an asynchronous system it is impossible to make sure that replicas of all stale configurations remove their private keys as it would require solving consensus [18]. However, as we show in this paper, it is possible to make sure that the configurations in which replicas do not remove their keys are never accessed by correct clients and are incapable of creating "proofs" for output values.

Finally, there is a subtle, and quite interesting "slow reader" attack. Suppose that a client hears from \textit{almost all} replicas in a quorum of the current configuration each holding a stale state, and not yet from the only correct replica in the quorum that has the up-to-date state. The client then falls asleep. Meanwhile, the configuration is superseded by a new one. As we do not make any assumptions about the correctness of replicas in stale configurations, the replica that has not yet responded can be compromised. Moreover, due to asynchrony, this replica can still retain its original private key. The replica can then pretend to be unaware of the current state. Therefore, the slow client might still be able to complete its request in the superseded configuration and return a stale state, which would violate the safety properties of the system. In Section 4, we give a detailed example of this attack and show that it can be addressed by an additional, "confirming" round-trip executed by the client.

Our contribution: Byzantine fault-tolerant reconfigurable services. We provide a systematic solution to each of the challenges described above and present a set of techniques for building reconfigurable services in asynchronous model with Byzantine faults of both clients and replicas. We consider a very strong adversary: any number of clients can be Byzantine and, as soon as some configuration is installed, no assumptions are made about the correctness of replicas in any of the prior configurations.

Moreover, in our quest for a simple solution for the Byzantine model, we devised a new approach to building asynchronous reconfigurable services by further exploring the connection between reconfiguration and lattice agreement [24, 27]. We believe that this approach can be effectively applied to crash fault-tolerant systems as well. As we discuss in Section 5.2, the proposed protocol has the time complexity that is optimal even for crash fault-tolerant systems.

Instead of trying to build a complex graph of configurations "on the fly" while simultaneously transferring the state between these configurations, we start by simply assuming that we are already given a \textit{linear history} (i.e., a sequence of configurations). We introduce the notion of a \textit{dynamic object} – an object that can transfer its own state between the configurations of a given finite linear history and serve meaningful user requests. We then provide dynamic implementations of several important object types, such as Lattice Agreement and Max-Register. We expect that other asynchronous static algorithms can be translated to the dynamic model using similar techniques.

Finally, we present a \textit{general transformation} that allows us to combine any dynamic object with two Dynamic Byzantine Lattice Agreement objects in such a way that together they constitute a single \textit{reconfigurable} object, which exports a general-purpose reconfiguration interface and supports all the operations of the original dynamic object.

This paper is a revised and extended version of a conference article [28].

Roadmap. The rest of the paper is organized as follows. We overview the model assumptions in Section 2 and define our principal abstractions in Section 3. In Section 4, we describe our implementation of Dynamic Byzantine Lattice Agreement. In Section 5, we show how to use it to implement reconfigurable objects out of dynamic ones. In Section 6, several possible implementations of access control are discussed. We discuss related work in Section 7 and conclude in Section 8.

The proof of correctness for our Dynamic Byzantine Lattice Agreement abstraction is delegated to Appendix A. Finally, as an application of our constructions, we provide an implementation of a dynamic Max-Register in Appendix B.

2 \textbf{SYSTEM MODEL}

\textbf{Processes and channels.} We consider a system of \textit{processes}. A process can be a \textit{replica} or a \textit{client}. Let \( \Phi \) and \( \Pi \) denote the (possibly infinite) sets of replicas and clients, resp., that potentially can take part in the computation. At any point in a given execution, a process can be in one of the four states: \textit{idle}, \textit{correct}, \textit{halted}, or \textit{Byzantine}. Initially, each process is \textit{idle}. An \textit{idle} process does not participate in the protocol. Once a process starts executing the protocol and as long as it does not execute the "halt" command and does not deviate from the prescribed protocol, it is considered \textit{correct}. A process is \textit{halted} if it executed the special "halt" command and stopped taking further steps. Finally, a process is \textit{Byzantine} if it
prematurely stops taking steps of the algorithm or takes steps that are not prescribed by it. A correct process can later halt or become Byzantine. However, the reverse is impossible: a halted or Byzantine process cannot become correct. We assume that a process that remains correct forever (we call such processes forecasts) does not prematurely stop taking steps of its algorithm.

We assume asynchronous reliable authenticated point-to-point links between each pair of processes [13]. If a forecast-correct process \( p \) sends a message \( m \) to a forecast-correct process \( q \), then \( q \) eventually delivers \( m \). Moreover, if a correct process \( q \) receives a message \( m \) from a process \( p \) at time \( t \), and \( p \) is correct at time \( t \), then \( p \) has indeed sent \( m \) to \( q \) before \( t \).

We assume that the adversary is computationally bounded so that it is unable to break the cryptographic techniques, such as digital signatures, forward security schemes [10] and one-way hash functions.

**Configuration lattice.** A join semi-lattice (or simply a lattice) is a tuple \((\mathcal{L}, \sqsubseteq)\), where \( \mathcal{L} \) is a set partially ordered by the binary relation \( \sqsubseteq \) such that for all elements \( x, y \in \mathcal{L} \), there exists the least upper bound for the set \( \{x, y\} \), i.e., the element \( z \in \mathcal{L} \) such that \( x, y \subseteq z \) and \( \forall w \in \mathcal{L} \) if \( x, y \subseteq w \), then \( z \subseteq w \). The least upper bound for the set \( \{x, y\} \) is denoted by \( x \sqcup y \). \( \sqcup \) is called the join operator. It is an associative, commutative, and idempotent binary operator on \( \mathcal{L} \). We write \( x \sqsubseteq y \) whenever \( x \subseteq y \) and \( x \neq y \). We say that \( x, y \in \mathcal{L} \) are comparable iff either \( x \sqsubseteq y \) or \( y \sqsubseteq x \).

For any (potentially infinite) set \( A \), \((2^A, \Subseteq)\) is a join semi-lattice, called the powerset lattice of \( A \). For all \( Z = (Z_1, Z_2) \in 2^A, Z_1 \subseteq Z_2 \subseteq Z_1 \cup Z_2 \). A configuration is an element of a join semi-lattice \((\mathcal{C}, \sqsubseteq)\). We assume that every configuration is associated with a finite set of replicas via a map \( \text{replicas} : \mathcal{C} \rightarrow 2^\Phi \), and a quorum system via a map \( \text{quorums} : \mathcal{C} \rightarrow 2^\Phi \), such that \( \forall C \in \mathcal{C} : \text{quorums}(C) \subseteq 2^{\text{replicas}(C)} \). Additionally, we assume that there is a map \( \text{height} : \mathcal{C} \rightarrow \mathbb{Z} \), such that \( \forall C \in \mathcal{C} : \text{height}(C) \geq 0 \) and \( \text{height}(C) \leq \text{height}(C') \), where \( \text{height}(C) < \text{height}(C') \). We say that a configuration \( C \) is higher (resp., lower) than a configuration \( D \) iff \( D \sqsubseteq C \) (resp. \( C \sqsubseteq D \)). Note that \( C \) is higher than \( D \) implies \( \text{height}(C) > \text{height}(D) \), but not vice versa.

We say that \( \text{quorums}(C) \) is a dissemination quorum system at time \( t \) iff every two sets \( \{a, b\} \) found within \( \text{quorums}(C) \) have at least one replica in common that is correct at time \( t \), and at least one quorum is available (all its replicas are correct) at time \( t \).

A natural (but not the only possible) way to define the lattice \( \mathcal{C} \) as follows: let \( \text{Updates} \) be \( \{+, -\} \times \Phi \), where tuple \((+, p)\) means “add replica \( p \)” and tuple \((-p)\) means “remove replica \( p \)”. Then \( \mathcal{C} \) is the powerset lattice \((\text{Updates} \sqsubseteq)\). The mappings \( \text{replicas} \), \( \text{quorums} \), and \( \text{height} \) are defined as follows: \( \text{replicas}(C) \subseteq \{s \in \Phi \mid (+, s) \in C \land (-, s) \notin C\} \). It is straightforward to verify that \( \text{quorums}(C) \) is a dissemination quorum system when strictly less than one third of replicas in \( \text{replicas}(C) \) are faulty. Note that, when this lattice is used for configurations, once a replica is removed from the system, it cannot be added again with the same identifier. In order to add such a replica back to the system, a new identifier must be used.

**Forward-secure digital signatures.** In a forward-secure digital signature scheme [10, 11, 16, 31], the public key of a process is fixed while the secret key can evolve. Each signature is associated with a timestamp. To generate a signature with timestamp \( t \), the signer uses secret key \( sk_t \). The signer can update its secret key and get \( sk_t \) from \( sk_{t+1} \) if \( t_1 < t_2 \leq T \). However “downgrading” the key to a lower timestamp is computationally infeasible. Thus, if the signer updates their secret key to some timestamp \( t \) and then removes the original secret key, it will not be able to sign new messages with a timestamp lower than \( t \), even if it later turns Byzantine.

For simplicity, we model a forward-secure signature scheme as an oracle which associates every process \( p \) with a timestamp \( st_p \) (initially, \( st_p = 0 \)). The oracle provides \( p \) with three operations:

1. \( \text{UpdateFSKey}(t) \) sets \( st_p \) to \( t \).
2. \( \text{FSSign}(m, t) \) returns a signature for message \( m \) and timestamp \( t \) if \( t \leq st_p \), otherwise it returns \( \bot \).
3. \( \text{FSVerify}(m, p, s, t) \) returns true iff \( s \) was generated by invoking \( \text{FSSign}(m, t) \) by process \( p \).

In our protocols, we use the height of the configuration as the timestamp. When a replica answers requests in configuration \( C \), it signs messages with timestamp \( \text{height}(C) \). When a higher configuration \( D \) is installed, the replica invokes \( \text{UpdateFSKey}(\text{height}(D)) \). This prevents the “I still work here” attack described in the introduction.

### 3 ABSTRACTIONS AND DEFINITIONS

In this section, we introduce principal abstractions of this paper (the access-control interface, Byzantine Lattice Agreement, Reconfigurable and Dynamic objects), state our quorum assumptions, and recall the definitions of broadcast primitives used in our algorithms.

#### 3.1 Access control and object composition

In our implementations and definitions, we parameterize some abstractions by boolean functions \( \text{VerifyInputValue}(v, \sigma) \) and \( \text{VerifyInputConfig}(C, \sigma) \), where \( \sigma \) is called a certificate. Moreover, some objects also export a boolean function \( \text{VerifyOutputValue}(v, \sigma) \), which lets anyone to verify that the value \( v \) was indeed produced by the object. This helps us to deal with Byzantine clients. In particular, it achieves three important goals.

First, the parameter \( \text{VerifyInputConfig} \) allows us to prevent Byzantine clients from reconfiguring the system in an undesirable way or flooding the system with excessively frequent reconfiguration requests. In Section 6, we propose three simple implementations of this functionality: each reconfiguration request must be signed by a quorum of replicas of some configuration or by a quorum of preconfigured administrators.

Second, the parameter \( \text{VerifyOutputValue} \) allows us to formally capture the application-specific notions of well-formed client requests and access control. For example, in a key-value storage system, each client can be permitted to modify only the key-value pairs that were created by this client. In this case, the certificate \( \sigma \) is just a digital signature of the client.
Finally, the exported function \texttt{VerifyOutputValue} allows us to compose several distributed objects in such a way that the output of one object is passed as input for another one. For example, in Section 5, one object (\texttt{HistLA}) operates exclusively on outputs of another object (\texttt{ConfLA}). We use the parameter function \texttt{VerifyInputChange} of \texttt{HistLA} and the exported function \texttt{VerifyOutputValue} of \texttt{ConfLA} to guarantee that a Byzantine client cannot send to \texttt{HistLA} a value that was not produced by \texttt{ConfLA}.

### 3.2 Byzantine Lattice Agreement abstraction

In this section we formally define \textit{Byzantine Lattice Agreement} abstraction (\texttt{BLA} for short), which serves as one of the main building blocks for constructing reconfigurable objects. Byzantine Lattice Agreement is an adaptation of Lattice Agreement [17] that can tolerate Byzantine failures of processes (both clients and replicas). It is parameterized by a join semi-lattice $L$, called the \textit{object lattice}, and a boolean function \texttt{VerifyInputChange} : $L \times \Sigma \rightarrow \{true, false\}$, where $\Sigma$ is a set of possible certificates. We say that $\sigma$ is a valid certificate for input value $v$ iff \texttt{VerifyInputChange}$(v, \sigma) = true$.

We say that $v \in L$ is a \textit{verifiable input value} in a given run iff at some point in time in that run, some process knows a certificate $\sigma$ that is valid for $v$, i.e., it maintains $v$ and a valid certificate $\sigma$ in its local memory. We require that the adversary is unable to invert \texttt{VerifyInputChange} by computing a valid certificate for a given value. This is the case, for example, when $\sigma$ must contain a set of unforgeable digital signatures.

The Byzantine Lattice Agreement abstraction exports one operation and one function:

- **Operation \texttt{Propose}(v, \sigma)** returns a response of the form $(w, \tau)$, where $v, w \in L$, $\sigma$ is a valid certificate for input value $v$, and $\tau$ is a certificate for output value $w$;
- **Function \texttt{VerifyOutputValue}(v, \sigma)** returns a boolean value.

Similarly to input values, we say that $\tau$ is a valid certificate for output value $w$ iff \texttt{VerifyOutputValue}(w, $\tau$) = true. We say that $w$ is a \textit{verifiable output value} in a given run iff at some point in that run, some process knows $\tau$ that is valid for $w$.

Implementations of Byzantine Lattice Agreement must satisfy the following properties:

- **BLA-Validity:** Every verifiable output value $w$ is a join of some set of verifiable input values;
- **BLA-Verifiability:** If \texttt{Propose}(\ldots) returns $(w, \tau)$ to a correct process, then \texttt{VerifyOutputValue}(w, $\tau$) = true;
- **BLA-Inclusion:** If $\texttt{Propose}(v, \sigma)$ returns $(w, \tau)$ to a correct process, then $v \subseteq w$;
- **BLA-Comparability:** All verifiable output values are comparable;
- **BLA-Liveness:** If the total number of verifiable input values is finite, every call to \texttt{Propose}(v, $\sigma$) by a forever-correct process eventually returns.

For the sake of simplicity, we only guarantee liveness when there are finitely many verifiable input values. This is sufficient for the purposes of reconfiguration, as it only guarantees liveness under the assumption that only finitely many valid reconfiguration calls are issued. In practice, this assumption boils down to providing liveness when not “too many” conflicting values are concurrently proposed. The abstraction that provides unconditional liveness is called \textit{Generalized Lattice Agreement} [17].

### 3.3 Reconfigurable objects

It is possible to define a \textit{reconfigurable} version of every static distributed object by enriching its interface and imposing some additional properties. In this section, we define the notion of a reconfigurable object in a very abstract way. By combining this definition with the definition of a Byzantine Lattice Agreement from Section 3.2, we obtain a formal definition of a Reconfigurable Byzantine Lattice Agreement. Similar combination can be performed with the definition of any static distributed object (e.g., with the definition of a Max-Register from Appendix B).

A reconfigurable object exports an operation \texttt{UpdateConfig}(C, $\sigma$), which can be used to reconfigure the system, and must be parameterized by a boolean function \texttt{VerifyInputChange} : $C \times \Sigma \rightarrow \{true, false\}$, where $\Sigma$ is a set of possible certificates. Similarly to verifiable input values, we say that $C \in C$ is a \textit{verifiable input configuration} in a given run iff at some point in that run, some process knows $\sigma$ such that \texttt{VerifyInputChange}(C, $\sigma$) = true.

We require the total number of verifiable input configurations to be finite in any given infinite execution of the protocol. In practice, this boils down to assuming sufficiently long periods of stability when no new verifiable input configurations appear. This requirement is imposed by all asynchronous reconfigurable storage systems [1, 3, 27, 35] we are aware of, and, in fact, can be shown to be necessary [34].

When a correct replica $r$ is ready to serve user requests in a configuration $C$, it triggers \texttt{InstalledConfig}(C). We then say that $r$ \textit{installs} configuration $C$. At any given moment in time, a configuration is called \textit{installed} if some correct replica has installed it and it is called \textit{superseded} if some higher configuration is installed.

Each reconfigurable object must satisfy the following properties:

- **Reconfiguration Validity:** Every installed configuration $C$ is a join of some set of verifiable input configurations. Moreover, all installed configurations are comparable;
- **Reconfiguration Liveness:** Every call to \texttt{UpdateConfig}(C, $\sigma$) by a forever-correct client eventually returns. Moreover, $C$ or a higher configuration will eventually be installed;
- **Installation Liveness:** If some configuration $C$ is installed by some correct replica, then $C$ or a higher configuration will eventually be installed by all correct replicas.

### 3.4 Dynamic objects

Reconfigurable objects are hard to build because they need to solve two problems at once. First, they need to order and combine concurrent reconfiguration requests. Second, the state of the object needs to be transferred across installed configurations. We decouple these two problems by introducing the notion of a \textit{dynamic} object. Dynamic objects solve the second problem while “outsourcing” the first one.
Before we formally define dynamic objects, let us first define the notion of a history. In Section 2, we introduced the configuration lattice $C$. A finite set $h \subseteq C$ is called a history iff all elements of $h$ are comparable (in other words, if they form a sequence). Let $\text{HighestConf}(h)$ be $C \in h$ such that $\forall C' \in h : C' \subseteq C$. $\text{HighestConf}(h)$ is well-defined, as the configurations in $h$ are totally ordered.

Dynamic objects must export an operation $\text{UpdateHistory}(h, \sigma)$ and must be parameterized by a boolean function $\text{VerifyHistory} : H \times \Sigma \rightarrow \{\text{true}, \text{false}\}$, where $H$ is the set of all histories and $\Sigma$ is the set of all possible certificates. We say that $h$ is a verifiable history in a given run if at some point in that run, some process knows $\sigma$ such that $\text{VerifyHistory}(h, \sigma) = \text{true}$. A configuration $C$ is called candidate iff it belongs to some verifiable history. Also, a candidate configuration $C$ is called active iff it is not superseded by a higher configuration.

As with verifiable input configurations, the total number of verifiable histories is required to be finite. Additionally, we require all verifiable histories to be related by containment (i.e., comparable w.r.t. $\subseteq$). Recall that a history is a totally ordered (w.r.t. $\subseteq$) set of configurations. Formally, if $\text{VerifyHistory}(h_1, \sigma_1) = \text{true}$ and $\text{VerifyHistory}(h_2, \sigma_2) = \text{true}$, then $h_1 \subseteq h_2$ or $h_2 \subseteq h_1$. We discuss how to build such histories in Section 5.

Similarly to reconfigurable objects, a dynamic object must have the $\text{InstalledConfig}(C)$ upcall. The object must satisfy the following properties:

- **Dynamic Validity**: Only a candidate configuration can be installed by a correct replica;
- **Dynamic Liveness**: Every call to $\text{UpdateHistory}(h, \sigma)$ by a forever-correct client eventually returns. Moreover, $\text{HighestConf}(h)$ or a higher configuration will eventually be installed;
- **Installation Liveness** (the same as for reconfigurable objects): If some configuration $C$ is installed by some correct replica, then $C$ or a higher configuration will eventually be installed by all correct replicas.

Note that Dynamic Validity implies that all installed configurations are comparable, since all verifiable histories are related by containment and all configurations within one history are comparable.

While reconfigurable objects provide general-purpose reconfiguration interface, dynamic objects are weaker, as they require an external service to build comparable verifiable histories. As the main contribution of this paper, we show how to build dynamic objects in a Byzantine environment and how to create reconfigurable objects using dynamic objects as building blocks. We argue that this technique is applicable to a large class of objects.

### 3.5 Quorum system assumptions

Most fault-tolerant implementations of distributed objects impose some requirements on the subsets of processes that can be faulty. We say that a configuration $C$ is correct at time $t$ iff $\text{repli}(C)$ is a dissemination quorum system at time $t$ (as defined in Section 2). Correctness of our implementations of dynamic objects relies on the assumption that active candidate configurations are correct. Once a configuration is superseded by a higher configuration, we make no further assumptions about it.

For reconfigurable objects we impose a slightly more conservative requirement: every combination of verifiable input configurations that is not yet superseded must be correct. Formally, we require:

#### Quorum availability

Let $C_1, \ldots, C_k$ be verifiable input configurations such that $C = C_1 \cup \cdots \cup C_k$ is not superseded at time $t$. Then we require quorum$(C)$ to be a dissemination quorum system at time $t$.

Correctness of our reconfigurable objects relies solely on correctness of the dynamic building blocks. Formally, when $k$ configurations are concurrently proposed, we require all possible combinations, i.e., $2^k - 1$ configurations, to be correct. However, in practice, at most $k$ of them will be chosen to be put in verifiable histories, and only those configurations will be accessed by correct processes. We impose a more conservative requirement because we do not know these configurations a priori.

### 3.6 Broadcast primitives

To make sure that no process is “left behind”, we assume that a variant of reliable broadcast primitive [13] is available. The primitive must ensure two properties:

1. If a forever-correct process $p$ broadcasts a message $m$, then $p$ eventually delivers $m$.
2. If some message $m$ is delivered by a forever-correct process, every forever-correct process eventually delivers $m$.

Note that we do not make any assumptions involving any processes that are not forever-correct. In practice such a primitive can be implemented by a gossip protocol [25]. This primitive is “global” in a sense that it is not bound to any particular configuration. In pseudocode we use “RB-Broadcast (...)” to denote a call to the “global” reliable broadcast.

Additionally, we assume a “local” uniform reliable broadcast primitive [13]. It has a stronger totality property: if some correct process $p$ delivered some message $m$, then every forever-correct process will eventually deliver $m$, even if $p$ later turns Byzantine. This primitive can be implemented in a static system, provided a quorum system. As we deal with dynamic systems, we associate every broadcast message with a fixed configuration and only guarantee these properties if the configuration is never superseded. Note that any static implementation of uniform reliable broadcast trivially guarantees this property. In pseudocode we use “URB-Broadcast (...)” in $C$ to denote a call to the “local” uniform reliable broadcast in configuration $C$.

### 4 Dynamic Byzantine Lattice Agreement

Dynamic Byzantine Lattice Agreement (DBLA for short) is the main building block in our construction of reconfigurable objects. Its specification is a combination of the specification of Byzantine Lattice Agreement (Section 3.2) and the specification of a dynamic object (Section 3.4). It is summarized in Algorithm 1. In this section, we provide the implementation of DBLA and analyze the time complexity of the solution. The proof of correctness is delegated to Appendix A.

As we mentioned earlier, we use forward-secure digital signatures to guarantee that superseded configurations cannot affect
Algorithm 1 DBLA object specification

Parameters:
1: Lattice of configurations C and the initial configuration Cinit \in C
2: The object lattice \mathcal{L} and the initial value Vinit \in \mathcal{L}
3: Boolean functions VerifyHistory(h, \sigma) and VerifyInputValue(v, \sigma)

Interface:
4: operation Propose(v, \sigma)
5: operation UpdateHistory(h, \sigma)
6: function VerifyOutputValue(v, \sigma)
7: upcall InstalledConfig(C)

Properties: BLA-Validity, BLA-Verifiability, BLA-Inclusion, BLA-Comparability, BLA-Liveness,
8: Dynamic Validity, Dynamic Liveness, Installation Liveness

Correct clients or forge certificates for output values. Ideally, before a new configuration C is installed (i.e., before a correct replica triggers InstalledConfig(C) upcall), we would like to make sure that the replicas of all candidate configurations lower than C invoke UpdateFSKey(height(C)). However, this would require the replica to know the set of all candidate configurations lower than C. Unambiguously agreeing on this set would require solving consensus, which is known to be impossible in a fault-prone asynchronous system [18].

Instead, we classify all candidate configurations in two categories: pivotal and tentative. A candidate configuration is called pivotal if it is the highest configuration in some verifiable history. Otherwise, it is called tentative. A nice property of pivotal configurations is that it is impossible to "skip" one in a verifiable history. Indeed, if \( C_1 = \text{HighestConf}(h_1) \) and \( C_2 = \text{HighestConf}(h_2) \) and \( C_1 \subseteq C_2 \), then, since all verifiable histories are related by containment, \( h_1 \subseteq h_2 \) and \( C_1 \in C_2 \). This allows us to make sure that, before a configuration C is installed, the replicas in all pivotal (and, possibly, some tentative) configurations lower than C update their keys.

In order to reconfigure a DBLA object, a correct client must use reliable broadcast to distribute the new verifiable history. Each correct process p maintains, locally, the largest (with respect to \( \subseteq \)) verifiable history it delivered so far through reliable broadcast. It is called the local history of process p and is denoted by \( \text{history}_p \). We use \( \text{Highest}_p \) to denote the most recent configuration in p’s local history (i.e., \( \text{Highest}_p = \text{HighestConf}(\text{history}_p) \)). Whenever a replica p updates \( \text{history}_r \), it invokes UpdateFSKey(height(Chighest_r)). Recall that if at least one forever-correct process delivers a message via reliable broadcast, every other forever-correct process will eventually deliver it as well.

Similarly, each process p keeps track of all verifiable input values it has seen \( \text{curVals}_p \subseteq \mathcal{L} \times \Sigma \), where \( \mathcal{L} \) is the object lattice and \( \Sigma \) is the set of all possible certificates. Sometimes, during the execution of the protocol, processes exchange these sets. Whenever a process p receives a message that contains a set of values with certificates \( \forall v \subseteq \mathcal{L} \times \Sigma \), it checks that the certificates are valid (i.e., \( \forall v (v, \sigma) \in vs : \text{VerifyInputValue}(v, \sigma) = \text{true} \)) and adds these values and certificates to \( \text{curVals}_p \).

4.1 Client implementation

The client’s protocol is simple. The pseudocode for it is presented in Algorithms 2 and 3. Note that we omit the subscript p in the pseudocode because each process can access only its own variables directly.

As we mentioned earlier, the operation UpdateHistory(h, \sigma) is implemented as RB-Broadcast (NewHistory, h, \sigma) (line 30). The rest of the reconfiguration process is handled by the replicas. The protocol for the operation Propose(v, \sigma) (lines 24–29) consists of two stages: proposing a value and confirming the result.

The first stage (proposing) mostly follows the implementation of lattice agreement by Faleiro et al. [17]. Client p repeatedly sends message \( \langle \text{Propose}, \text{curVals}_p, \text{seqNum}_p, C \rangle \) to all replicas in \text{replicas}(C), where \( \text{Propose} \) is the message descriptor, \( C = \text{Chighest}_p \), and \( \text{seqNum}_p \) is a sequence number used by the client to match sent messages with replies.

After sending these messages to \text{replicas}(C), the client waits for responses of the form \( \langle \text{ProposeResp}, vs, sig, sn \rangle \), where \( \text{ProposeResp} \) is the message descriptor, \( vs \) is the set of all verifiable input values known to the replica with valid certificates (including those sent by the client), \( sig \) is a forward-secure signature with timestamp \( \text{height}(C) \), and \( sn \) is the sequence number as in the message from the client.

During the first stage, one of the following cases can take place: (1) the client learns about some new verifiable input values from one of the \text{ProposeResp} messages; (2) the client updates its local history (by delivering it through reliable broadcast); and (3) the client receives a quorum of valid replies with the same set of verifiable input values. In the latter case, the client proceeds to the second stage. In the first two cases, the client simply restarts the operation. Recall that, according to the BLA-Liveness property, termination of client requests is only guaranteed when the number of verifiable input values is finite. Additionally, the number of verifiable histories is assumed to be finite. Hence, the number of restarts will also be finite. This is the main intuition behind the liveness of the client’s protocol.

The example in Figure 1 illustrates how the first stage of the algorithm ensures the comparability of the results when no reconfiguration is involved. In this example, clients p and q concurrently propose values \{1\} and \{2\}, respectively, from the lattice \( \mathcal{L} = 2^N \).
Algorithm 2 DBLA: code for client $p$ (part 1)

Parameters:
9: Lattice of configurations $C$ and the initial configuration $C_{init}$
10: The object lattice $L$ and the initial value $V_{init}$
11: Boolean functions $\text{VerifyHistory}(h, \sigma)$ and $\text{VerifyInputValue}(a, \sigma)$

Global variables:
12: $\text{history} \subseteq C$, initially $\{C_{init}\}$
13: $\sigma_{\text{history}} \in \Sigma$, initially $\perp$
14: $\text{curVals} \subseteq L \times \Sigma$, initially $\{(V_{init}, L)\}$
15: $\text{status} \in \{\text{inactive, proposing, confirming}\}$, initially inactive
16: $\text{seqNum} \in \mathbb{Z}$, initially 0
17: $\text{acks}_1$, initially $\emptyset$
18: $\text{acks}_2$, initially $\emptyset$

Auxiliary functions:
19: $\text{HighestConf}(h)$
20: $\text{ContainsQuorum}(\text{acks}, C)$
21: $\text{JoinAll}(\text{vs})$
22: $\text{VerifyInputValue}(a, \sigma)$
23: $\text{FSVerify}(m, r, s, t)$

operation $\text{Propose}(a, \sigma)$
24: $\text{Refine} \{(a, \sigma)\}$
25: $\text{wait for } \text{ContainsQuorum}(\text{acks}_2, \text{HighestConf}(\text{history}))$
26: $\text{status} \leftarrow \text{inactive}$
27: $\text{return } \text{JoinAll}(\text{curVals}, \sigma)$

operation $\text{UpdateHistory}(h, \sigma)$
28: $\text{VerifyOutputValue}(a, \sigma)$
29: if $\sigma = \perp$ then return $a = V_{init}$
30: let $\sigma = (\text{curVals}, \text{history}, \sigma_{\text{history}}, \text{acks}_1, \text{acks}_2)$
31: let $C = \text{HighestConf}(h)$
32: $\text{return } \text{JoinAll}(\text{vs}) = a \land \text{VerifyHistory}(h, \sigma_h)$
33: \quad $= \text{ContainsQuorum}(\text{proposeAcks}, C) \land \text{ContainsQuorum}(\text{confirmAcks}, C)$
34: \quad $\land \forall (r, s) \in \text{proposeAcks} : \text{FSVerify}(\langle \text{ProposeResp}, \text{vs}, r, s, \text{height}(C) \rangle)$
35: \quad $\land \forall (r, s) \in \text{confirmAcks} : \text{FSVerify}(\langle \text{ConfirmResp}, \text{proposeAcks}, r, s, \text{height}(C) \rangle)$

Client $p$ successfully returns the proposed value $\{1\}$ while client $q$ is forced to refine its proposal and return the combined value $\{1, 2\}$. The quorum intersection prevents the clients from returning incomparable values (e.g., $\times\{1\}$ and $\{2\}$).

In the second (confirming) stage of the protocol, the client simply sends the acknowledgments it has collected in the first stage to the replicas of the same configuration. The client then waits for a quorum of replicas to reply with a forward-secure signature with timestamp $\text{height}(C)$.

The example in Figure 2 illustrates how reconfiguration can interfere with an ongoing $\text{Propose}$ operation in what we call the “slow reader” attack, and how the second stage of the protocol prevents a safety violation. Imagine that a correct client completed the $\text{Propose}$ operation and received value $\{2\}$ before client $p$ started the execution. As a result, all correct replicas in quorum $\{r_1, r_2, r_3\}$ store value $\{2\}$. Then client $p$ executes $\text{Propose}(\{1\}, \sigma)$, where $\sigma$ is a valid certificate for input value $\{1\}$. Due to the $\text{BLA-Comparability}$, $\text{BLA-Inclusion}$, and $\text{BLA-Validity}$ properties of Byzantine Lattice Agreement, the only valid output value for client $p$ is $\{1, 2\}$ (assuming that there are no other verifiable input values). The client successfully reaches replicas $r_2$ and $r_3$ before the reconfiguration. Neither $r_1$ nor $r_2$ tells the client about the input value $\{2\}$. $r_2$ is outdated and $r_3$ is Byzantine. The message from $p$ to $r_1$ is delayed. Meanwhile, a new configuration is installed, and all replicas of the original configuration become Byzantine. When the message from $p$ finally reaches $r_1$, the replica is already Byzantine and it can pretend that it has not seen any verifiable input values other than $\{1\}$. The client then finishes the first stage of the protocol with value $\{1\}$. Returning this value from from the $\text{Propose}$ operation would violate $\text{BLA-Comparability}$. Luckily, the second stage of the protocol prevents the safety violation. Since replicas $r_2$ and $r_3$ updated their private keys during the reconfiguration, they are unable to send the signed confirmations with timestamp $\text{height}(C)$ to the client. Hence, the client will not
be able to complete the operation in configuration \( C \) and will wait until it receives a new verifiable history via reliable broadcast and will restart the operation in a higher configuration.

The certificate for the output value \( v \in \mathcal{L} \) produced by the \textsc{Propose} protocol in a configuration \( C \) consists of: (1) the set of verifiable input values (with certificates for them) from the first stage of the algorithm (the join of all these values must be equal to \( v \); (2) a verifiable history (with a certificate for it) that confirms that

\begin{algorithm}
\begin{algorithmic}
\Procedure{Refine}{\( vs \)}
\State \( acks_1 \leftarrow \emptyset; \; acks_2 \leftarrow \emptyset \)
\State \( curVals \leftarrow curVals \cup vs \)
\State \( seqNum \leftarrow seqNum + 1 \)
\State \( status \leftarrow \text{proposing} \)
\State let \( C = \text{HighestConf}(\text{history}) \)
\State \textbf{send} \( (\text{Propose}, \; curVals, \; seqNum, \; C) \) to replicas(\( C \))
\EndProcedure
\Endalgorithm
\end{algorithm}
C is pivotal (i.e., for which C is the highest configuration); (3) the quorum of signatures from the first stage of the algorithm; and (4) the quorum of signatures from the second stage of the algorithm. Intuitively, the only way for a Byzantine client to obtain such a certificate is to benignly follow the PROPOSE protocol.

It is important that only pivotal configurations can produce valid certificates because non-pivotal (tentative) configurations may contain fully compromised quorums with non-updated private keys for the forward-secure digital signature scheme.

4.2 Replica implementation

The pseudocode for the replicas is presented in Algorithms 4 and 5.

Each replica r maintains, locally, its current configuration (denoted by Ccurr, r) and the last configuration installed by this replica (denoted by Cinstr). Cinst ⊆ Ccurr ⊆ Chighest, r. Intuitively, Ccurr, r = C means that replica r knows that there is no need to transfer state from configurations lower than C, either because r already performed the state transfer from those configurations, or because it knows that sufficiently many other replicas did. Cinstr, r = C means that the replica knows that sufficiently many replicas in C have up-to-date states, and that configuration C is ready to serve user requests.

As we saw earlier, each client message is associated with some configuration C. The replica only processes the message when C = Cinstr, r = Ccurr, r = Chighest, r. If C § Chighest, r, the replica simply ignores the message. Due to the properties of reliable broadcast, the client will eventually learn about Chighest, and will repeat its request there (or in an even higher configuration). If Cinstr, r ⊆ C and Chighest, r ⊆ C, the replica waits until C is installed before processing the message. Finally, if C is incomparable with Cinstr, or Chighest, r, then, since all candidate configurations are required to be comparable, the message is sent by a Byzantine client and the replica should ignore it.

When a correct replica r receives a Propose message (line 72), it adds the newly learned verifiable input values to curVals, r and sends curVals, r back to the client with a forward-secure signature with timestamp height(C). When a correct replica receives a Confirm message (line 79), it simply signs the set of acknowledgments in it with a forward-secure signature with timestamp height(C) and sends the signature to the client.

A very important part of the replica’s implementation is the state transfer protocol. The pseudocode for it is presented in Algorithm 5. Let Cnext, r be the highest configuration in history, such that r ∈ replicas(Cnext, r). Whenever Ccurr, r ≠ Cnext, r, the replica tries to “move” to Cnext, r by reading the current state from all configurations between Ccurr, r and Cnext, r one by one in ascending order (line 90). In order to read the current state from configuration C ⊆ Cnext, r, replica r sends message (UpdateRead, seqNum, r, C) to all replicas in replicas(C). In response, each replica r1 ∈ replicas(C) sends curVals, r1, r to r in an UpdateReadResp message (line 103). However, r1 replies only after its private key is updated to a timestamp larger than height(C) (line 102). We maintain the invariant (line 100) that for every correct replica q, the timestamp stq is always equal to height.(Chighest).q.

If r receives a quorum of replies from the replicas of C, there are two distinct cases:

- C is still active at the moment when r receives the last acknowledgment. In this case, the quorum intersection property still holds for C, and replica r can be sure that (1) if some Propose operation has completed in configuration C or reached the second stage with some set of verifiable input values vs, then vs ⊆ curVals, r; and (2) if some Propose operation has not yet reached the second stage, it will not be able to complete in configuration C (it will have to retry in a higher configuration, see the example in Figure 2).
- C is already superseded by the time r receives the last acknowledgment. This means that some configuration higher than C is installed, and the state from configuration C was already transferred to that higher configuration. We refer to Appendix A for more formal proofs.

If a configuration C is superseded before r receives enough replies, it may happen that r will never be able to collect a quorum of replies from C. However, in this case, r will eventually discover that some higher configuration is installed, and it will update Ccurr, r (line 109). The waiting on line 92 will terminate due to the first part of the condition (C ⊆ Ccurr).

When a correct replica completes transferring the state to some configuration C, it notifies other replicas about it by broadcasting message UpdateComplete in configuration C (line 95). A correct replica installs a configuration C if it receives such messages from a quorum of replicas in C (line 106). Because we want our protocol to satisfy the Installation Liveness property (if one correct replica installs a configuration, every forever-correct replica must eventually install this or a higher configuration), the UpdateComplete messages are distributed through the uniform reliable broadcast primitive that we introduced in Section 3.6.

4.3 Time complexity

In our analysis we assume that the time complexity of the reliable broadcast primitive, which we use to disseminate verifiable histories, is constant. With this assumption, it is easy to see that the worst-case time complexity of our DBLA implementation is O(m + k), where m is the number of verifiable input values and k is the size of the largest verifiable history. Indeed, the time complexity is proportional to the number of calls to Refine. There are only two reasons why a client may call Refine: either it learns about a new verifiable input value (line 54), which may happen at most m times or it learns about a new verifiable history (line 62), which may happen at most k times.

As for the complexity of the state transfer protocol, it is linear in the size of the largest verifiable history k. Because a replica advances the Curr variable at the end of state transfer (line 94), each of the k candidate configurations is accessed at most once by each replica, and in each configuration, our state transfer protocol makes a constant number of steps.

4.4 Implementing other dynamic objects

While we do not provide any general approach for building dynamic objects, we expect that most asynchronous Byzantine fault-tolerant static algorithms can be adapted to the dynamic case by applying the same set of techniques. These techniques include our state transfer protocol (relying on forward-secure signatures), the use of
To illustrate this, in Appendix B, we present the dynamic version of Max-Register \[5\]. We also discuss the dynamic version of the Access Control abstraction in Section 6.

\[\text{Algorithm 4 DBLA: code for replica } r \text{ (part 1)}\]

\begin{verbatim}
Parameters: \(C, L, Cinit, Vinit, \text{VERIFYINPUTVALUE}(h, \sigma), \text{and VERIFYINPUTVALUE}(c, \sigma)\) (see Algorithm 2)

Global variables:
63: \(\text{history} \subseteq C\), initially \(\{\text{Cinit}\}\)
64: \(\text{curVals} \subseteq L \times \Sigma\), initially \(\{\langle Vinit, L \rangle\}\)
65: \(\text{Curr} \in C\), initially \(\text{Cinit}\)
66: \(\text{Cinst} \in C\), initially \(\text{Cinit}\)
67: \(\text{seqNum} \in \mathbb{Z}\), initially 0
68: \(\text{inStateTransfer} \in \{\text{true, false}\}\), initially false

Auxiliary functions:
69: \(\text{HighestConf, ContainsQuorum, JoinAll, VERIFYINPUTVALUES}\) (see Algorithm 2).
70: \(\text{FSSign}(\text{message, timestamp})\)
71: \(\text{UpdateFSKey}(t)\)

72: \textbf{upon receive (Propose, vs, sn, C) from client } c
73: \hspace{1em} \textbf{wait for } C = \text{HighestConf}(\text{history}) \cup \text{Cinst} \hspace{1em} \triangleright \text{known verifiable input values with proofs}
74: \hspace{1em} \textbf{if } C = \text{HighestConf}(\text{history}) \wedge \text{VERIFYINPUTVALUES}(\text{vs} \setminus \text{curVals}) \hspace{1em} \triangleright \text{current configuration}
75: \hspace{1em} \text{curVals} \leftarrow \text{curVals} \cup \text{vs}
76: \hspace{1em} \text{let } \text{sig} = \text{FSSign}(\langle \text{ProposeResp, curVals, height(C)} \rangle)
77: \hspace{1em} \text{send } \langle \text{ProposeResp, curVals, sig, sn} \rangle \text{ to } c \hspace{1em} \triangleright \text{updated signing timestamp (see Section 2)}
78: \hspace{1em} \text{else ignore the message}

79: \textbf{upon receive (Confirm, proposeAcks, sn, C) from client } c
80: \hspace{1em} \textbf{wait for } C = \text{Cinst} \cup \text{HighestConf}(\text{history}) \cup \text{Cinst} \hspace{1em} \triangleright \text{installed configuration}
81: \hspace{1em} \textbf{if } C = \text{HighestConf}(\text{history}) \hspace{1em} \triangleright \text{used to match requests with responses}
82: \hspace{1em} \text{let } \text{sig} = \text{FSSign}(\langle \text{ConfirmResp, proposeAcks, height(C)} \rangle)
83: \hspace{1em} \text{send } \langle \text{ConfirmResp, sig, sn} \rangle \text{ to } c \hspace{1em} \triangleright \text{produces a forward-secure signature (see Section 2)}
84: \hspace{1em} \text{else ignore the message}
\end{verbatim}

Figure 3: The structure of dependencies in our implementation of a reconfigurable object. The arrow from object \(A\) to object \(B\) represents a dependency of object \(A\) on object \(B\). The label next to the arrow reflects the nature of this dependency.

5 IMPLEMENTING RECONFIGURABLE OBJECTS

While dynamic objects are important building blocks, they are not particularly useful by themselves because they require an external source of comparable verifiable histories. In this section, we show how to combine several dynamic objects to obtain a single reconfigurable object. Similar to dynamic objects, the specification of a reconfigurable object can be obtained as a combination of the specification of a static object with the specification of an abstract reconfigurable object from Section 3.3. In particular, compared to static objects, reconfigurable objects have one more operation – UpdateConfig\((C, \sigma)\), must be parameterized by a boolean function VerifyINPUTCONFIG\((C, \sigma)\), and must satisfy Reconfiguration Validity, Reconfiguration Liveness, and Installation Liveness.

We build a reconfigurable object by combining three dynamic ones. The first one is the dynamic object that executes clients’ operations (let us call it \(\text{DObj}\)). For example, in order to implement a reconfigurable version of Byzantine Lattice Agreement, one needs to take a dynamic version of Byzantine Lattice Agreement as \(\text{DObj}\). Similarly, in order to implement a reconfigurable version of Max-Register \[5\], one needs to take a dynamic version of Max-Register as \(\text{DObj}\) (see Appendix B). The two remaining objects are used to build verifiable histories: \(\text{ConfLA}\) is a DBLA operating on the configuration lattice \(C\), and \(\text{HistLA}\) is a DBLA operating on the powerset lattice \(2^C\). The relationships between the three dynamic objects are depicted in Figure 3.

an additional round-trip to prevent the “slow reader” attack, and the structure of our cryptographic proofs ensuring that tentative configurations cannot create valid certificates for output values. To illustrate this, in Appendix B, we present the dynamic version of Max-Register \[5\]. We also discuss the dynamic version of the Access Control abstraction in Section 6.
Algorithm 5 DBLA: code for replica $r$ (part 2)

```plaintext
85: upon $C_{curr} \neq \text{HIGHESTConf}(\{C \in \text{history} \mid r \in \text{replicas}(C)\}) \land \text{not inStateTransfer}$
86: let $C_{next} = \text{HIGHESTConf}(\{C \in \text{history} \mid r \in \text{replicas}(C)\})$
87: let $S = \{C \in \text{history} \mid C_{curr} \subseteq C \subseteq C_{next}\}$
88: $\text{inStateTransfer} \leftarrow \text{true}$
89: $\text{seqNum} \leftarrow \text{seqNum} + 1$
90: for each $C \in S$ in ascending order do
91:   send $(\text{UpdateRead}, \text{seqNum}, C)$ to replicas$(C)$
92: wait for $(C \subseteq C_{curr}) \lor \text{responses from any } Q \in \text{quorums}(C)$ with s.n. seqNum
93: if $C_{curr} \subseteq C_{next}$ then
94:   $C_{curr} \leftarrow C_{next}$
95: $\text{URB-Broadcast}\langle \text{UpdateComplete} \rangle$ in $C_{next}$
96: $\text{inStateTransfer} \leftarrow \text{false}$

97: upon RB-deliver $\langle \text{NewHistory}, h, \sigma \rangle$ from any sender
98: if $\text{VerifyHistory}(h, \sigma) \land \text{history} \subseteq h$ then
99:   history $\leftarrow h$
100: $\text{UpdateFSKey}(\text{height}(\text{HIGHESTConf}(\text{history})))$

101: upon receive $\langle \text{UpdateRead}, sn, C \rangle$ from replica $r'$
102: wait for $(C \subseteq \text{HIGHESTConf}(\text{history}))$
103: send $(\text{UpdateReadResp}, \text{curVals}, sn)$ to $r'$

104: upon receive $\langle \text{UpdateReadResp}, vs, sn \rangle$ from replica $r'$
105: if $\text{VerifyInputValues}(\text{vs} \setminus \text{curVals})$ then $\text{curVals} \leftarrow \text{curVals} \cup \text{vs}$
106: upon URB-deliver $\langle \text{UpdateComplete} \rangle$ in $C$ from quorum $Q \in \text{quorums}(C)$
107: wait for $C \in \text{history}$
108: if $C_{inst} \subseteq C$ then
109:   if $C_{curr} \subseteq C$ then $C_{curr} \leftarrow C$
110:   $C_{inst} \leftarrow C$
111: trigger upcall $\text{INSTALLEDCONFIG}(C)$
112: if $r \notin \text{replicas}(C)$ then $\text{halt}$
```

The pseudocode is presented in Algorithm 6. All data operations are performed directly on $DObj$. To update a configuration, the client first submits its proposal to $ConfLA$ and then submits the result as a singleton set to $HistLA$. Due to the BLA-Comparability property, all verifiable output values produced by $ConfLA$ are comparable, and any combination of them would create a well-formed history as defined in Section 3.4. Moreover, the verifiable output values of $HistLA$ are related by containment, and, therefore, can be used as verifiable histories in dynamic objects. We use them to reconfigure all three dynamic objects (lines 123–125).

Cryptographic keys. In Algorithm 6, we use several dynamic objects. We assume that correct replicas have separate public/private key pairs for each dynamic object. This prevents replay attacks across objects and allows each dynamic object to manage its keys separately. We discuss how to avoid this assumption later in this section.

5.1 Proof of correctness

In the following two lemmas we show that we use the dynamic objects ($ConfLA$, $HistLA$, and $DObj$) correctly, i.e., all requirements imposed on verifiable histories are satisfied.

**Lemma 5.1.** All histories passed to the dynamic objects by correct processes (lines 123–125) are verifiable with $\text{VerifyHistory}$ (line 133).

**Proof.** Follows from the BLA-Verifiability property of $HistLA$. □

**Lemma 5.2.** All histories verifiable with $\text{VerifyHistory}$ (line 133) are (1) well-formed (that is, consist of comparable configurations) and (2) related by containment. Moreover, (3) in any given infinite execution, there is only a finite number of histories verifiable with $\text{VerifyHistory}$.

**Proof.** (1) follows from the BLA-Comparability property of $ConfLA$, the BLA-Validity property of $HistLA$, and the definition of $HistLA.\text{VERIFYINPUTVALUE}$ (line 129).

(2) follows directly from the BLA-Comparability property of $HistLA$.

(3) follows from the requirement of finite number of verifiable input configurations and the BLA-Validity property of $ConfLA$ and $HistLA$. Only a finite number of configurations can be formed by $ConfLA$ out of a finite number of verifiable input configurations, and only a finite number of histories can be formed by $HistLA$ out of the configurations produced by $ConfLA$. □
Algorithm 6 Reconfigurable object

```
\begin{itemize}
\item Common code
\end{itemize}

Parameters:
1. Lattice of configurations $C$ and the initial configuration $C_{init}$
2. Boolean function $\text{VerifyInputConfig}(C, \sigma)$
3. Dynamic object $DObj$, which we want to make reconfigurable

Shared objects:
1. $DObj$
2. $ConfLA$
3. $HistLA$

\item Code for client $p$

1. Data operations are performed directly on $DObj$.

\item operation $\text{UpdateConfig}(C, \sigma)$

1. let $\langle D, \sigma_D \rangle = \text{ConfLA}.\text{Propropose}(C, \sigma)$
2. let $\langle h, \sigma_h \rangle = \text{HistLA}.\text{Propose}(\langle D, \sigma_D \rangle)$
3. $DObj.\text{UpdateHistory}(h, \sigma_h)$
4. $ConfLA.\text{UpdateHistory}(h, \sigma_h)$
5. $HistLA.\text{UpdateHistory}(h, \sigma_h)$

\item Code for replica $r$

1. upon receive upcall $\text{InstalledConfig}(C)$ from all $DObj$, $ConfLA$, and $HistLA$
2. trigger upcall $\text{InstalledConfig}(C)$

\item Parameters specification

1. function $\text{ConfLA}.\text{VerifyInputValue}(v, \sigma) = \text{VerifyInputConfig}(v, \sigma)$
2. function $\text{HistLA}.\text{VerifyInputValue}(v, \sigma)$
3. if $v$ is not a set of 1 element then return false
4. let $\langle C \rangle = v$
5. return $\text{ConfLA}.\text{VerifyOutputValue}(C, \sigma)$

\item All dynamic objects are parameterized with the same $\text{VerifyHistory}$ function.

\item function $\text{VerifyHistory}(h, \sigma) = \text{HistLA}.\text{VerifyOutputValue}(h, \sigma)$
```

Theorem 5.3 (Transformation safety). Our implementation satisfies the Reconfiguration Validity property of a reconfigurable object. That is, (1) every visited configuration $C$ is a join of some set of verifiable input configurations; and (2) all installed configurations are comparable.

Proof. (1) follows from the BLA-Validity property of $ConfLA$ and $HistLA$ and the Dynamic Validity property of the underlying dynamic objects. (2) follows directly from the Dynamic Validity property of the underlying dynamic objects.

Theorem 5.4 (Transformation liveness). Our implementation satisfies the liveness properties of a reconfigurable object: Reconfiguration Liveness and Installation Liveness.

Proof. Reconfiguration Liveness follows from the BLA-Liveness property of $ConfLA$ and $HistLA$ and the Dynamic Liveness property of the underlying dynamic objects. Installation Liveness follows from line 126 of the implementation and the Installation Liveness of the underlying dynamic objects.

5.2 Discussion

Time complexity: By accessing $ConfLA$ and then $HistLA$, we minimize the number of configurations that should be accessed for a consistent configuration shift. Indeed, due to the BLA-Validity property of $ConfLA$ and $HistLA$, when $k$ reconfiguration requests are executed concurrently, at most $k$ new verifiable histories will be created and the total number of candidate configurations will not exceed $k+1$ (including the initial configuration). As a result, only $O(k)$ configurations are accessed for state transfer, similar to [20] and [35]. In contrast, in DynaStore [2], a client might have to access up to $\Omega(\min\{mk, 2^k\})$ configurations, where $m$ is the number of concurrent data operations.

Additionally, every reconfiguration request involves two invocations of DBLA-Proposel. The worst-case latency of our DBLA-Propose implementation is $O(k + m)$, where $m$ is the number of verifiable input values. In this case, $m = k$. Hence, the worst-case latency of a reconfiguration request is linear with respect to the number of verifiable input configurations, which is known to be optimal even for crash fault-tolerant systems [35].

Bootstrapping. The relationship between lattice agreement and reconfiguration has been studied before [24, 27]. In particular, as

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The analogy for the number of verifiable input configurations in crash fault-tolerant systems is the number of reconfiguration requests. In the context of Byzantine fault-tolerant systems, we have to talk about the number of verifiable input configurations instead because a single Byzantine client can simulate an infinite sequence of requests.
shown in [24], lattice agreement can be used to build comparable configurations. We take a step further and use two separate instances of lattice agreement: one to build comparable configurations (ConfLA) and the other to build histories out of them (HistLA). These two LA objects can then be used to reconfigure a single dynamic object (DObj).

However, this raises a question: how to reconfigure the lattice agreement objects themselves? We found the answer in the idea that is sometimes referred to as “bootstrapping”. We use the lattice agreement objects to reconfigure themselves and at least one other object. This implies that the lattice agreement objects share the configurations with DObj. The most natural implementation is that the code for all three dynamic objects (ConfLA, HistLA, and DObj) will be executed by the same set of replicas.

Bootstrapping is a dangerous technique because, if applied badly, it can lead to infinite recursion. However, we structured our solution in such a way that there is no recursion at all: the client first makes normal requests to ConfLA and HistLA and then uses the resulting history to reconfigure all dynamic objects, as if this history was obtained by the client from the outside of the system. It is important to note that liveness of the call HistLA.VERIFYOUTPUTVALUE(h, σ) is not affected by reconfiguration: the function simply checks some digital signatures and is guaranteed to always terminate given enough processing time.

Shared parts. All implementations of dynamic objects presented in this paper have a similar structure. For example, they all share the same state transfer implementation (see Algorithm 5). However, we do not deny the possibility that other implementations of other dynamic objects might have very different implementations. Therefore, in our transformation we use DObj as a “black box” and do not make any assumptions about its implementation. Moreover, for simplicity, we use the two DBLA objects as “black boxes” as well. In fact, ConfLA and HistLA may have different implementations and the transformation will still work as long as they satisfy the specification from Section 3. However, this comes at a cost.

In particular, if implemented naively, a single reconfigurable object will run several independent state transfer protocols, and a single correct replica will have several private/public key pairs (as mentioned earlier in this section). But if, as in this paper, all dynamic objects have similar implementations of their state transfer protocols, this can be done more efficiently by combining all state transfer protocols into one, which would need to transfer the states of all dynamic objects and make sure that the superseded configurations are harmless.

6 ACCESS CONTROL

Our implementation of reconfigurable objects relies on the parameter function VERIFYINPUTCONFIG. Moreover, if we apply our transformation from Section 5 to our implementation of DBLA from Section 4, the resulting reconfigurable object will rely on the parameter function VERIFYINPUTVALUE. The implementation of these parameters is highly application-specific. For example, in a storage system, it is reasonable to only allow requests that modify some data if the are accompanied by a digital signature of the owner of the data. For the sake of completeness, in this section, we present three generic implementations for these parameter functions. We believe that each of these implementations is suitable for some applications.

In order to do this, we introduce the Access Control object. It exports one operation and one function:

- Operation REQUESTCERT(σ) returns a certificate σ, which can be verified with VERIFYCERT(σ, σ), or the special value ⊥, indicating that the permission was denied;
- Function VERIFYCERT(σ, σ) returns a boolean value.

The implementation of Access Control must satisfy the following property:

- Certificate Verifiability: If REQUESTCERT(σ) returned σ to a correct process, then VERIFYCERT(σ, σ) = true.

In Sections 6.1–6.3, we present three different implementations of the dynamic version of the Access Control object, and in Section 6.4, we show how to use it in order to implement the parameter functions VERIFYINPUTVALUE and VERIFYINPUTCONFIG.

6.1 Trusted administrators

A naive yet common approach to dynamic systems is to have a pre-configured trusted administrator, who signs the reconfiguration requests. However, if the administrator’s private key is lost, the system might lose liveness, and if it is compromised, the system might lose even safety. A more viable approach is to have n administrators and to require b + 1 of them to sign every certificate, for some n and b such that 0 ≤ b < n. In this case, the system will “survive” up to b keys being compromised and up to n – (b + 1) keys being lost.

6.2 Sanity-check approach

One of the simplest implementations of access control in a static system is to require at least b + 1 replicas to sign each certificate, where b is the maximal possible number of Byzantine replicas, sometimes called the resilience threshold. The correct replicas can perform some application-specific sanity checks before approving requests.

The key property of this approach is that it guarantees that each valid certificate is signed by at least one correct replica. In many cases, this is sufficient to guarantee resilience both against the Sybil attacks [15] and against attempts to flood the system with reconfiguration requests. The correct replicas can check the identities of the new participants and refuse to sign excessively frequent requests.

In dynamic asynchronous systems, just providing b+1 signatures is not sufficient. Despite the use of forward-secure signatures, in a superseded pivotal configuration there might be significantly more than b Byzantine replicas with their private keys not removed (in fact, at least 2b). The straightforward way to implement this policy in a dynamic system is to add the confirming phase, as in our implementation of Dynamic Byzantine Lattice Agreement (see Section 4), after collecting b+1 signed approvals. The confirming phase guarantees that, during the execution of the first phase, the configuration was active. The state transfer protocol should be the same as for DBLA with the exception that no actual state is being transferred. The only goal of the state transfer protocol in this case...
Algorithm 7 Vote-Based Dynamic Access Control

\[ \text{operation RequestCert}(v) \]
\[ \text{let } C = \text{HighestConf} \text{(history)} \]
\[ \text{seqNum }\leftarrow \text{seqNum }+ 1 \]

\[ \text{Phase one: request} \]
\[ \text{send } \langle \text{Request, v, seqNum, C} \rangle \text{ to replicas(C)} \]
\[ \text{wait for} \text{(HighestConf \((\text{history}) \neq C \) \lor \text{(received enough Yes-votes with valid signatures)}} \]
\[ \text{if HighestConf \((\text{history}) \neq C \text{ then } \text{restart the operation (goto line 135)}} \]
\[ \text{let } \text{acks}_1 = \{ \text{Yes-votes received on line 140} \} \]

\[ \text{Phase two: confirm} \]
\[ \text{send } \langle \text{Confirm, acks}_1, \text{seqNum, C} \rangle \text{ to replicas(C)} \]
\[ \text{wait for} \text{(HighestConf \((\text{history}) \neq C \) \lor \text{(a quorum of replies with valid signatures)}} \]
\[ \text{if HighestConf \((\text{history}) \neq C \text{ then } \text{restart the operation (goto line 135)}} \]
\[ \text{let } \text{acks}_2 = \{ \text{acknowledgments received on line 145} \} \]

\[ \text{Return certificate} \]
\[ \text{return } \langle \text{history, } \sigma_{\text{history}}, \text{acks}_1, \text{acks}_2 \rangle \]

\[ \text{Code for replica r} \]
\[ \text{upon receive } \langle \text{Request, v, sn, C} \rangle \text{ from client c} \]
\[ \text{wait for } C = \text{Cinst }\lor C \in \text{HighestConf \text{(history)}} \]
\[ \text{if } C = \text{HighestConf \text{(history)}} \text{ then} \]
\[ \text{if VoteYes(a) then send } \langle \text{Yes, FSignature(Yes, v, c), height(C)}, \text{sn} \rangle \text{ to c} \]
\[ \text{else send } \langle \text{No, sn} \rangle \text{ to c} \]

\[ \text{upon receive } \langle \text{Confirm, acks, sn, C} \rangle \text{ from client c} \]
\[ \text{wait for } C \in \text{history} \]
\[ \text{if } C = \text{HighestConf \text{(history)}} \text{ then} \]
\[ \text{let } \text{sig} = \text{FSignature(ConfirmResp, acks), height(C)} \]
\[ \text{send } \langle \text{ConfirmResp, sig, sn} \rangle \text{ to c} \]

is to make sure that the replicas update their private keys before a new configuration is installed.

This and the following approach can be generally described as “vote-based” access control policies. The pseudocode for their dynamic implementation is presented in Algorithm 7.

6.3 Quorum-based approach (“on-chain governance”)

A more powerful strategy in a static system is to require a quorum of replicas to sign each certificate. An important property of this implementation is that it can detect and prevent conflicting requests. More formally, suppose that there are values \( v_1 \) and \( v_2 \), for which the following two properties should hold:

- Both are acceptable: \( \text{RequestCert}(v_i) \) should not return \( \perp \) unless \( \text{RequestCert}(v_i) \) was invoked in the same execution, where \( j \neq i \).
- At most one may be accepted: if some process knows \( \sigma_i \) such that \( \text{VerifyCert}(v_i, \sigma_i) \) then no process should know \( \sigma_j \) such that \( \text{VerifyCert}(v_j, \sigma_j) \).

Note that it is possible that neither \( v_1 \) nor \( v_2 \) is accepted by the Access Control if the requests are made concurrently. To guarantee that exactly one certificate is issued, we would need to implement consensus, which is impossible in asynchronous model [18]. If a correct replica has signed a certificate for value \( v_1 \), it should store this fact in persistent memory and refuse signing \( v_j \) if requested. Due to the quorum intersection property, this guarantees the “at most one” semantic in a static system.

This approach can be implemented in a dynamic system using the pseudocode from Algorithm 7 and the state transfer protocol from our DBLA implementation (see Algorithm 5).

Using the dynamic version of this approach to certifying reconfiguration requests allows us to capture the notion of what is sometimes called “on-chain governance”. The idea is that the participants of the system (in our case, the owners of the replicas) decide which actions or updates to allow by the means of voting. Every decision needs a quorum of signed votes to be considered valid and no two conflicting decisions can be made.
Early proposals of (actively) reconfigurable storage systems tolerating process crashes, such as RAMBO [23] and reconfigurable Paxos [29], used consensus (and, thus, assumed certain level of synchrony) to ensure that the clients agree on the evolution of configurations. DynaStore [2] was the first asynchronous reconfigurable storage: clients propose incremental additions or removals to the system configuration. As the proposals commute, the processes can resolve their disagreements without involving consensus.

The parsimonious speculative snapshot task [20] resolves conflicts between concurrent configuration updates in a storage system using instances of commit-adopt [19]. The worst-case time complexity, in the number of message delays, of reconfiguration was later reduced from $O(nm)$ [20] to $O(n + m)$ [35], where $n$ is the number of concurrently proposed configuration updates and $m$ is the number of concurrent data operations. This coincides with the time complexity of our solution.

SmartMerge [24] made an important step forward by treating reconfiguration as an instance of abstract lattice agreement [17]. However, the algorithm assumes an external (reliable) lattice agreement service which makes the system not fully reconfigurable.

FreeStore [3] describes an algorithm for reconfigurable storage that can be seen as a composition of a reconfiguration protocol and a read-write protocol. Reconfiguration is based on the view generator abstraction, which encapsulates the form of agreement required to reconcile concurrently proposed reconfiguration requests (essentially, very similar to lattice agreement). The use of view generators helps in optimizing latency, similar to our use of dynamic lattice agreement objects.

The recently proposed reconfigurable lattice-agreement abstraction [27] enables reconfigurable versions of a large class of objects and constructions, including state-based CRDTs [33], atomic snapshot, max-register, conflict detector and commit-adopt. Configurations are treated here in an abstract way, as elements of a configuration lattice, encapsulating replica sets and quorum assumptions. We believe that the reconfiguration service we introduced in this paper can be used to derive Byzantine fault-tolerant reconfigurable implementations of objects in the class.

Byzantine quorum systems [30] introduce abstractions for ensuring availability and consistency of shared data in asynchronous systems with Byzantine faults. In particular, a dissemination quorum system ensures that every two quorums have a correct process in common and that at least one quorum only contains correct processes.

Dynamic Byzantine quorum systems [4] appear to be the first attempt to implement a form of active reconfiguration in a Byzantine fault-tolerant data service running on a static set of replicas, where clients can raise or lower the resilience threshold. Dynamic Byzantine storage [32] allows a trusted administrator to issue ordered reconfiguration calls that might also change the set of replicas. The administrator is also responsible for generating new private keys for the replicas in each new configuration to anticipate the "I still work here" attack [1]. In this paper, we propose an implementation of a Byzantine fault-tolerant reconfiguration service that does not rely on this assumption.

Forward-secure signature schemes [10, 11, 14, 16, 31] were originally designed to mitigate the consequences of key exposure: if the private key of an agent is compromised, signatures made prior to

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**Figure 4:** Two possible ways to integrate the Access Control abstraction with other types of objects. An arrow from an object $A$ to another object $B$ marked with VIC, (resp., VIC or VH) indicates that $A$.$\text{VERIFYINPUTVALUE}$ (resp., $A$.$\text{VERIFYINPUTCONFIG}$ or $A$.$\text{VERIFYHISTORY}$) is implemented using $B$.$\text{VERIFYOUTPUTVALUE}$ or $B$.$\text{VERIFYCERT}$.

### 6.4 Combining Access Control with other objects

There are at least two possible ways to combine the Access Control abstraction with a reconfigurable object in a practical system.

The simplest, and, perhaps, the most practical approach is to embed two instances of Dynamic Access Control directly into the structure of a reconfigurable object, as shown in Figure 4a. In this case, the replicas that execute the code for the Access Control are the same replicas as the replicas that execute the code of other dynamic objects in the implementation of this reconfigurable object.

Alternatively, one can apply the transformation from Section 5 to the dynamic Access Control implementation described in this section to obtain a reconfigurable version of the Access Control abstraction. It then can be combined with any other reconfigurable object in a structure depicted in Figure 4b. In this case, the replicas of ConfRAC produce verifiable input configurations for themselves and for two other objects.
the exposure (i.e., with smaller timestamps) can still be trusted. In this paper, a novel application of forward-secure digital signatures is proposed: timestamps are associated with configurations. Before a new configuration is installed, the protocol ensures that sufficiently many correct processes update their private keys in prior configurations. This approach prevents the “I still work here” and “slow reader” attacks. Unlike previously proposed solutions [32], it was designed for similar applications.

We would like to mention a few possible directions. Most of them are dedicated to reducing the communication cost of the protocol.

First, the proofs in our protocol include the full local history of a process. Moreover, this history comes with its own proof, which also usually contains a history, and so on. If implemented naively, the size of one proof in bytes will be at least quadratic with respect to the number of distinct candidate configurations, which is not necessary. The first observation is that these histories will be related by containment. So, in fact, they can be compressed just to the size of the largest one, which is linear. But we can go further and say that, in fact, in a practical implementation, the processes almost never should actually send full histories to each other because every process maintains its local history and all histories with proofs are already disseminated via the reliable broadcast primitive. When one process wants to send some history to some other process, it can just send a cryptographic hash of this history. The other process can check if it already has this history and, if not, ask the sender to only send the missing parts, instead of the whole history.

Second, a naive implementation of our DBLA protocol would send ever-growing sets of verifiable input values, which is also not necessary. The processes should just limit themselves to sending diffs between what they know and what they think the recipient knows.

Third, almost every proof in our systems contains signatures from a quorum of replicas. This adds another linear factor to the communication cost. However, it can be significantly reduced by the use of forward-secure multi-signatures, such as Pixel [16], which was designed for similar applications.

Finally, we use a suboptimal implementation of lattice agreement as the foundation for our DBLA protocol. Perhaps, we could benefit from adapting a more efficient crash fault-tolerant asynchronous solution [36].

Open questions. We would like to mention two relevant directions for further research.

First, with regard to active reconfiguration, it would be interesting to devise algorithms that efficiently adapt to “small” configuration changes, while still supporting the option of completely changing the set of replicas in a single reconfiguration request. In this paper, we allow each reconfiguration request to completely change the set of replicas, which leads to an expensive quorum-to-quorum communication pattern. This seems unnecessary for reconfiguration requests involving only slight changes to the set of replicas.

Second, with regard to Byzantine faults, it would be interesting to consider models with a “weaker” adversary. In this paper, we assumed a very strong model of the adversary: no assumptions are made about correctness of replicas in superseded configurations. This “pessimistic” approach leads to more complicated and expensive protocols.

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A PROOF OF CORRECTNESS OF DBLA

A.1 Safety

Recall that a configuration is called candidate iff it appears in some verifiable history. The following lemma gathers some obvious yet very useful statements about candidate configurations.

**Lemma A.1 (Candidate configurations).**

1. There is a finite number of candidate configurations.
2. All candidate configurations are comparable with $\sqsubseteq$.

**Proof.** The total number of verifiable histories is required to be finite, and each history is finite, hence (1). All verifiable histories are required to be related by containment and all configurations within one history are required to be comparable, hence (2). □

Recall that a configuration is called pivotal if it is the last configuration in some verifiable history. Non-pivotal candidate configurations are called tentative. Intuitively, the next lemma states that in the rest of the proofs we can almost always consider only pivotal configurations. Tentative configurations are both harmless and useless.

**Lemma A.2 (Tentative configurations).**

1. No correct client will ever make a request to a tentative configuration.
2. Tentative configurations cannot be installed.
3. A correct process will never invoke FSVERIFY with timestamp $\text{height}(C)$ for any tentative configuration $C$.
4. A correct replica will never broadcast any message via the uniform reliable broadcast primitive in a tentative configuration.

**Proof.** Follows directly from the algorithm. Both clients and replicas only operate on configurations that were obtained by invoking the function $\text{HighestConf}(h)$ on some verifiable configuration. □

The next lemma states that correct processes cannot “miss” any pivotal configurations in their local histories. This is crucial for the correctness of the state transfer protocol.

**Lemma A.3.** If $C \sqsubseteq \text{HighestConf}(h)$, where $C$ is a pivotal configuration and $h$ is the local history of a correct process, then $C \in h$.

**Proof.** Follows directly from the definition of a pivotal configuration and the requirement that all verifiable histories are related by containment (see Section 3.4). □

Recall that a configuration is called superseded iff some higher configuration is installed (see Section 3.3). A configuration is installed iff some correct replica has triggered the $\text{InstalledConfig}$ upcall (line 111). For this, the correct replica must receive a quorum of $\text{UpdateComplete}$ messages via the uniform reliable broadcast primitive (line 106).

**Theorem A.4 (Dynamic Validity).** Our implementation of DBLA satisfies Dynamic Validity. I.e., only a candidate configuration can be installed.

**Proof.** Follows directly from the implementation. A correct replica will not install a configuration until it is in the replica’s local history (line 107). □

In our algorithm, it is possible for a configuration to be installed after it was superseded. Imagine that a quorum of replicas broadcast $\text{UpdateComplete}$ messages in some configuration $C$ which is not yet installed. After that, before any replica delivers those messages, a higher configuration is installed, making $C$ superseded. It is possible that some correct replica $r \in \text{replicas}(C)$ that does not yet know that a higher configuration is installed, will deliver the broadcast messages and trigger the upcall $\text{InstalledConfig}(C)$ (line 111).

Let us call the configurations that were installed while being active (i.e., not superseded) “properly installed”. We will use this definition to prove next few lemmas.

**Lemma A.5.** The lowest properly installed configuration higher than configuration $C$ is the first installed configuration higher than $C$ in the real-time order.

**Proof.** Let $N$ be the lowest properly installed configuration higher than $C$. If some configuration higher than $N$ were installed earlier, then $N$ would not be properly installed (by the definition of a properly installed configuration). If some configuration between $C$ and $N$ were installed earlier, then $N$ would not be the lowest. □

The following lemma stipulates that our state transfer protocol makes the superseded pivotal configurations “harmless” by leveraging a forward-secure signature scheme.

**Lemma A.6 (Key update).** If a pivotal configuration $C$ is superseded, then no quorum of replicas in that configuration is capable of signing messages with timestamp $\text{height}(C)$, i.e., $\exists Q \in \text{quorums}(C) \text{ s.t. } \forall r \in Q : \text{st}_r \leq \text{height}(C)$.

**Proof.** Let $N$ be the lowest properly installed configuration higher than $C$. Let us consider the moment when $N$ was installed. By the algorithm, all correct replicas in some quorum $Q_N \in \text{quorums}(N)$ had to broadcast $\text{UpdateComplete}$ messages before $N$ was installed (line 106). Since $N$ was not yet superseded at that moment, there was at least one correct replica $r_N \in Q_N$.

By Lemma A.3, $C$ was in $r_N$’s local history whenever it performed state transfer to any configuration higher than $C$. By the protocol, a correct replica only advances its $\text{Curr}$ variable after executing the state transfer protocol (line 94) or right before installing a configuration (line 109). Since no configurations between $C$ and $N$ were yet installed, $r_N$ had to pass through $C$ in its state transfer protocol and to receive $\text{UpdateReadResp}$ messages from some quorum $Q_C \in \text{quorums}(C)$ (line 92).

Recall that correct replicas update their private keys whenever they learn about a higher configuration (line 100) and that they will only reply to message ($\text{UpdateRead}$, $sn_C$, $C$) once $C$ is not the highest configuration in their local histories (line 102). This means that all correct replicas in $Q_C$ actually had to update their private keys before $N$ was installed, and, hence, before $C$ was superseded. By the quorum intersection property, this means that in each quorum in $C$ at least one replica updated its private key to a timestamp higher than $\text{height}(C)$ and will not be capable of signing messages with timestamp $\text{height}(C)$ even if it becomes Byzantine. □

Note that in a tentative configuration there might be arbitrarily many Byzantine replicas that have not updated their private keys.
This is inevitable in asynchronous systems: forcing the replicas in tentative configurations to update their private keys would require solving consensus. This does not affect correct processes because,
as shown in Lemma A.2, tentative configurations are harmless. However, it is important to remember this when designing new dynamic protocols.

The following lemma implies that the state is correctly transferred between configurations.

**Lemma A.7 (State transfer correctness).**
If \( \sigma = \langle vs, h, \sigma_p, \text{proposeAtkl} \rangle \) is a valid proof for \( v \), then for each active configuration \( D \) such that \( \text{HighestConf}(h) \subseteq D \), there is a quorum \( Q_D \in \text{quorums}(D) \) such that for each correct replica \( r \in Q_D \), \( \text{vs} \subseteq \text{curVals}_r \).

**Proof.** Let \( C = \text{HighestConf}(h) \). We proceed by induction on the sequence of all properly installed configurations higher than \( C \). Let \( C \) denote this sequence by \( \mathcal{C} \). By the definition of a properly installed configuration, these are precisely the configurations that we consider in the statement of the lemma.

Let \( N \) be the lowest configuration in \( \mathcal{C} \). Let \( Q_C \in \text{quorums}(C) \) be a quorum of replicas whose signatures are in \( \text{proposeAtkl} \). Consider the moment of installation of \( N \). There must be a quorum \( Q_N \in \text{quorums}(N) \) in which all correct replicas broadcast \( \text{UpdateComplete}(N) \) before the moment of installation. For each correct replica \( r_N \in Q_N \), \( r_N \) passed with its state transfer protocol through configuration \( C \) and received \( \text{UpdateReadResp} \) messages from some quorum of replicas in \( C \). Note that at that moment configuration \( C \) was not yet superseded. By the quorum intersection property, there is at least one correct replica \( r_C \in Q_C \) that sent an \( \text{UpdateReadResp} \) message to \( r_N \) before \( Q_N \) was installed. Let the \( \text{UpdateReadResp} \) message contain \( \text{curVals}_r \) of \( r_C \) that must have contained a set of values that includes all values from \( C \). This proves the base case of the induction.

Let us consider any configuration \( D \in \mathcal{C} \) such that \( N \subseteq D \). Let \( M \) be the highest configuration in \( \mathcal{C} \) such that \( M \subseteq D \) (in other words, the closest to \( D \) in \( \mathcal{C} \)). Assume that the statement holds for \( M \), i.e., while \( M \) was active, there were a quorum \( Q_M \in \text{quorums}(M) \) such that for each correct replica \( r_M \in Q_M \), \( \text{vs} \subseteq \text{curVals}_r \). Similarly to the base case, let us consider a quorum \( Q_D \in \text{quorums}(D) \) such that every correct replica in \( Q_D \) reliably broadcast \( \text{UpdateComplete}(D) \) before \( D \) was installed. For each correct replica \( r_D \in Q_D \), by the quorum intersection property, there is at least one correct replica in \( Q_D \) that sent an \( \text{UpdateReadResp} \) message to \( r_D \). This replica attached its \( \text{curVals} \) to the message, which contained \( \text{vs} \). This proves the induction step and completes the proof. \( \square \)

The next lemma states that if two output values were produced in the same configuration, they are comparable. In a static system it could be proven by simply referring to the quorum intersection property. In a dynamic Byzantine system, however, to use the quorum intersection, we need to prove that the configuration was active during the whole period when the clients were exchanging data with the replicas. In other words, we need to prove that the "slow reader" attack is impossible. Luckily, we have the second stage of our algorithm designed for this sole purpose.

**Lemma A.8 (BLA-Comparability in one configuration).**
If \( \sigma_1 = \langle vs_1, h_1, \sigma_{h_1}, \text{proposeAtkl}_1, \text{confirmAtkl}_1 \rangle \) is a valid proof for output value \( v_1 \), and \( \sigma_2 = \langle vs_2, h_2, \sigma_{h_2}, \text{proposeAtkl}_2, \text{confirmAtkl}_2 \rangle \) is a valid proof for output value \( v_2 \), and \( \text{HighestConf}(h_1) = \text{HighestConf}(h_2) \), then \( v_1 \) and \( v_2 \) are comparable.

**Proof.** Let \( C = \text{HighestConf}(h_1) = \text{HighestConf}(h_2) \). By definition, the fact that \( \sigma \) is a valid proof for \( v \) implies that \( \text{VerifyOutputValue}(v, \sigma) = \text{true} \) (line 32). By the implementation, \( h_1 \) and \( h_2 \) are verifiable histories (line 36). Therefore, \( C \) is a pivotal configuration.

The set \( \text{confirmAtkl}_1 \) contains signatures from a quorum of replicas of configuration \( C \), with timestamp \( \text{height}(C) \). Each of these signatures had to be produced after each of the signatures in \( \text{proposeAtkl}_1 \) because they sign the message \( \langle \text{ConfirmAtkl}, \text{proposeAtkl}_1 \rangle \) (line 82). Combining this with the statement of Lemma A.6 (Key Update), it follows that at the moment when the last signature in the set \( \text{proposeAtkl}_1 \) was created, the configuration \( C \) was active (otherwise it would be impossible to gather \( \text{confirmAtkl}_1 \)). We can apply the same argument to the sets \( \text{proposeAtkl}_2 \) and \( \text{confirmAtkl}_2 \).

It follows that there are quorums \( Q_1, Q_2 \in \text{quorums}(C) \) and a moment in time \( t \) such that: (1) \( C \) is not superseded at time \( t \); (2) all correct replicas in \( Q_1 \) signed message \( \langle \text{ProposeAtkl}, v_1 \rangle \) before \( t \); and (3) all correct replica in \( Q_2 \) signed message \( \langle \text{ProposeAtkl}, v_2 \rangle \) before \( t \). Since \( C \) is not superseded at time \( t \), there must be a correct replica in \( Q_1 \cap Q_2 \) (due to quorum intersection), which signed both \( \langle \text{ProposeAtkl}, v_1 \rangle \) and \( \langle \text{ProposeAtkl}, v_2 \rangle \) (line 76). Since correct replicas only sign \( \langle \text{ProposeAtkl}, v \rangle \) messages with comparable sets of values\(^6\), \( v_1 \) and \( v_2 \) are comparable, i.e., either \( v_1 \subseteq v_2 \) or \( v_2 \subseteq v_1 \). Hence, \( v_1 = \text{JOINALL}(v_1) \) and \( v_2 = \text{JOINALL}(v_2) \) are comparable. \( \square \)

Finally, let us combine the two previous lemmas to prove the BLA-Comparability property of our DBLA implementation.

**Theorem A.9 (BLA-Comparability).** Our implementation of DBLA satisfies the BLA-Comparability property: That is, all verifiable output values are comparable.

**Proof.** Let \( \sigma_1 = \langle vs_1, h_1, \sigma_{h_1}, \text{proposeAtkl}_1, \text{confirmAtkl}_1 \rangle \) be a valid proof for output value \( v_1 \), and \( \sigma_2 = \langle vs_2, h_2, \sigma_{h_2}, \text{proposeAtkl}_2, \text{confirmAtkl}_2 \rangle \) be a valid proof for output value \( v_2 \). Also, let \( C_1 = \text{HighestConf}(h_1) \) and \( C_2 = \text{HighestConf}(h_2) \). Since \( h_1 \) and \( h_2 \) are verifiable histories (line 36), both \( C_1 \) and \( C_2 \) are pivotal by definition.

If \( C_1 = C_2 \), \( v_1 \) and \( v_2 \) are comparable by Lemma A.8.

Consider the case when \( C_1 \neq C_2 \). Without loss of generality, assume that \( C_1 \subseteq C_2 \). Let \( Q_1 \in \text{quorums}(C_2) \) be a quorum of replicas whose signatures are in \( \text{proposeAtkl}_2 \). Let \( t \) be the moment when first correct replica signed \( \langle \text{ProposeAtkl}, v_2 \rangle \). Correct replicas only start processing user requests in a configuration when this

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\(^6\)Indeed, set \( \text{curVals} \) at each correct replica can only grow, and the replicas only sign messages with the same set of verifiable input values as \( \text{curVals} \) (see lines 75–76).
configuration is installed (line 73). Therefore, by Lemma A.7, at time \( t \) there was a quorum of replicas \( Q_2 \in \text{quorums}(C_2) \) such that for every correct replica in \( Q_2 \), \( v_{S_1} \subseteq \text{curVals} \). By the quorum intersection property, there must be at least one correct replica in \( Q_1 \cap Q_2 \). Hence, \( v_{S_1} \subseteq v_{S_2} \) and \( \text{JoinAll}(v_{S_1}) \subseteq \text{JoinAll}(v_{S_2}) \). □

**Theorem A.10 (DBLA safety).** Our implementation satisfies the safety properties of DBLA: BLA-Validity, BLA-Verifiability, BLA-Inclusion, BLA-Comparability, and Dynamic Validity.

**Proof.**

- BLA-Validity follows directly from the implementation: a correct client collects verifiable input values and joins them before returning from Propose (line 29);
- BLA-Verifiability follows directly from how correct replicas form and check the proofs for output values (lines 28 and 34–39);
- BLA-Inclusion follows from the fact that the set \( \text{curVals} \) of a correct client only grows (line 42);
- BLA-Comparability follows from Theorem A.9;
- Finally, Dynamic Validity follows from Theorem A.4.

□

A.2 Liveness

**Lemma A.11 (History Convergence).** Local histories of all correct processes will eventually become identical.

**Proof.** Let \( p \) and \( q \) be any two forever-correct processes\(^7\). Suppose, for contradiction, that the local histories of \( p \) and \( q \) have diverged at some point and will never converge again. Recall that correct processes only adopt verifiable histories, and that we require the total number of verifiable histories to be finite. Therefore, there is some history \( h_p \), which is the the largest history ever adopted by \( p \), and some history \( h_q \) which is the the largest history ever adopted by \( q \). Since all verifiable histories are required to be related by containment, and we assume that \( h_p \neq h_q \), one of them must be a subset of the other. Without loss of generality, suppose that \( h_p \subset h_q \). Since \( q \) had to deliver \( h_q \) through reliable broadcast (unless \( h_q \) is the initial history) and \( q \) remains correct forever, \( p \) will eventually deliver \( h_q \) as well, and will adopt it. Hence, \( h_p \) is not the largest history ever adopted by \( p \). A contradiction. □

Next, we introduce an important definition, which we will use throughout the rest of the proofs.

**Definition A.12 (Maximal installed configuration).** In a given infinite execution, a maximal installed configuration is a configuration that eventually becomes installed and never becomes superseded.

**Lemma A.13 (Cmax existence).** In any infinite execution there is a unique maximal installed configuration.

**Proof.** By Lemma A.1 (Candidate configurations) and Theorem A.4 (Dynamic Validity), the total number of installed configurations is finite and they are comparable. Hence, we can choose a unique maximum among them, which is never superseded by definition. □

Let us denote the (unique) maximal installed configuration by \( C_{\text{max}} \).

**Lemma A.14 (Cmax installation).** The maximal installed configuration will eventually be installed by all correct replicas.

**Proof.** Since \( C_{\text{max}} \) is installed, by definition, at some point some correct replica has triggered upcall \( \text{INSTALLEDCONF}(C_{\text{max}}) \) (line 111). This, in turn, means that this replica delivered a quorum of \( \text{UpdateComplete} \) messages via the uniform reliable broadcast in \( C_{\text{max}} \) when it was correct. Therefore, even if this replica later becomes Byzantine, by definition of the uniform reliable broadcast, either \( C_{\text{max}} \) will become superseded (which is impossible), or every correct replica will eventually deliver the same set of \( \text{UpdateComplete} \) messages and install \( C_{\text{max}} \). □

**Lemma A.15 (State transfer progress).** State transfer (lines 85–96) executed by a forever-correct replica always terminates.

**Proof.** Let \( r \) be a correct replica executing state transfer. By Lemma A.1, the total number of candidate configurations is finite. Therefore, it is enough to prove that there is no such configuration that \( r \) will wait for replies from a quorum of that configuration indefinitely (line 92). Suppose, for contradiction, that there is such configuration \( C \).

If \( C \subset C_{\text{max}} \), then, by Lemma A.14, \( r \) will eventually install \( C_{\text{max}} \), and \( C_{\text{curr}} \) will become not lower than \( C_{\text{max}} \) (line 109). Hence, \( r \) will terminate from waiting through the first condition \( (C \subset C_{\text{curr}}) \). A contradiction.

Otherwise, if \( C_{\text{max}} \subset C \), then, by the definition of \( C_{\text{max}} \), \( C \) will never be superseded. Since \( r \) remains correct forever, by Lemma A.11 (History Convergence), every correct replica will eventually have \( C \) in its local history. Since we assume reliable links between processes (see Section 2), every correct replica in \( \text{replicas}(C) \) will eventually receive \( r \)'s \( \text{UpdateRead} \) message and will send a reply, which \( r \) will receive (line 103). Hence, the waiting on line 92 will eventually terminate through the second condition \( (r \) will receive responses from some \( Q \in \text{quorums}(C) \) with the correct sequence number) . A contradiction. □

Intuitively, the following lemma states that \( C_{\text{max}} \) is, in some sense, the “final” configuration. After some point every correct process will operate exclusively on \( C_{\text{max}} \). No correct process will know about any configuration higher than \( C_{\text{max}} \) or “care” about any configuration lower than \( C_{\text{max}} \).

**Lemma A.16.** \( C_{\text{max}} \) will eventually become the highest configuration in the local history of each correct process.

**Proof.** By Lemma A.11 (History Convergence), the local histories of all correct processes will eventually converge to the same history \( h \). Let \( D = \text{HIGHESTConf}(h) \). Since \( C_{\text{max}} \) is installed and never superseded, it cannot be higher than \( D \) (at least one correct replica will always have \( C_{\text{max}} \) in its local history).

Suppose, for contradiction, that \( C_{\text{max}} \subset D \). In this case, \( D \) is never superseded, which means that there is a quorum \( Q_D \in \text{quorums}(D) \) that consists entirely of forever-correct processes. By

\(^7\)If either \( p \) or \( q \) eventually halts or becomes Byzantine, their local histories are not required to converge.
Lemma A.11 (History Convergence), all replicas in \( Q_D \) will eventually have \( D \) in their local histories and will try to perform state transfer to it. By Lemma A.15, they will eventually succeed and install \( D \) — a contradiction with the definition of \( C_{max} \).

Theorem A.17 (BLA-Liveness). Our implementation of DBLA satisfies the BLA-Liveness property: if the total number of verifiable input values is finite, every call to Propose\((a, \sigma)\) by a forever-correct process eventually returns.

Proof. Let \( p \) be a forever-correct client that invoked Propose\((a, \sigma)\). By Lemma A.16, \( C_{max} \) will eventually become the highest configuration in the local history of \( p \). If the client’s request will not terminate by the time it learns about \( C_{max} \), the client will call Refine\((\emptyset)\) after it (line 62). By Lemma A.14, \( C_{max} \) will eventually be installed by all correct replicas. Since it will never be superseded, there will be a quorum of forever-correct replicas. Thus, every round of messages from the client will eventually be responded to by a quorum of correct replicas.

Since the total number of verifiable input values is finite, the client will call Refine only a finite number of times (line 54). After the last call to Refine, the client will inevitably receive acknowledgments from a quorum of replicas, and will proceed to sending Confirm messages (line 50). Again, since there is an available quorum of correct replicas that installed \( C_{max} \), the client will eventually receive enough acknowledgments and will complete the operation (line 27).

Theorem A.18 (DBLA liveness). Our implementation satisfies the liveness properties of DBLA: BLA-Liveness, Dynamic Liveness, and Installation Liveness.

Proof. BLA-Liveness follows from Theorem A.17. Dynamic Liveness and Installation Liveness follow directly from Lemmas A.16 and A.14 respectively.

B MAX REGISTER

Our methodology of constructing dynamic and reconfigurable objects is not limited to lattice agreement. In this section, we show how to create an atomic Byzantine fault-tolerant Max-Register in dynamic setting.

is a distributed object that has two operations: Read\((a, \sigma)\) and Write\((a, \sigma)\) and must be parametrized by a boolean function VerifyInputValue\((a, \sigma)\). As before, we say that \( \sigma \) is a valid certificate for input value \( a \) iff VerifyInputValue\((a, \sigma) = true \) and that value \( a \) is a verifiable input value iff some process knows \( \sigma \) such that VerifyInputValue\((a, \sigma) = true \). We assume that correct clients invoke Write\((a, \sigma)\) only if VerifyInputValue\((a, \sigma) = true \). We do not make any assumptions on the number of verifiable input values for this abstraction (i.e., it can be infinite).

The Max-Register object satisfies the following three properties:

- **MR-Validity**: if Read() returns value \( v \) to a correct process, then \( v \) is verifiable input value;
- **MR-Atomicity**: if some correct process \( p \) completed Write\((a, \sigma)\) or received \( v \) from Read() strictly before some correct process \( q \) invoked Read(), then the value returned to \( q \) must be greater than or equal to \( v \);
- **MR-Liveness**: every call to Read() and Write\((a, \sigma)\) by a forever-correct process eventually returns.

For simplicity, unlike Byzantine Lattice Agreement, our Max-Register does not provide the VerifyOutputValue\((a, \sigma)\) function.

B.1 Dynamic Max-Register implementation

In this section we present our implementation of the dynamic version of the Max-Register abstraction (Dynamic Max-Register or DMR for short). Overall, the “application” part of the implementation is very similar to the classical ABD algorithm [6], and the “dynamic” part of the implementation is almost the same as in DBLA.

Client implementation. From the client’s perspective, the two main procedures are Get\((a, \sigma)\) and Set\((a, \sigma)\) (not to be confused with the Read and Write operations). Set\((a, \sigma)\) (lines 175–180) is used to store the value on a quorum of replicas of the most recent configuration. It returns true if it manages to receive signed acknowledgments from a quorum of some configuration. Forward-secure signatures are used to prevent the “I still work here” attack. Since Set does not try to read any information from the replicas, it is not susceptible to the “slow reader” attack. Get\((a, \sigma)\) (lines 181–187) is very similar to Set\((a, \sigma)\) and is used to request information from a quorum of replicas of the most recent configuration. Since we do not provide the VerifyOutputValue\((a, \sigma)\) function, the replies from replicas are not signed (line 199). Therefore, Get\((a, \sigma)\) is susceptible to both the “I still work here” and “slow reader” attack when used alone. Later in this section we discuss how the invocation of Set\((a, \sigma)\) right after Get\((a, \sigma)\) (line 167) allows us to avoid these issues.

Operation Write\((a, \sigma)\) (lines 170–172) is used by correct clients to store values in the register. It simply performs repeated calls to Write\((a, \sigma)\) and is used to request information from the replicas. Retries are safe because, as in lattice agreement, write requests to a max-register are idempotent. Since we assume the total number of verifiable histories to be finite, only a finite number of retries is possible.

Operation Read\((a, \sigma)\) (lines 164–169) is used to request the current value from the register, and it consists of repeated calls to both Get\((a, \sigma)\) and Set\((a, \sigma)\). The call to Get\((a, \sigma)\) is simply used to query information from the replicas. The call to Set\((a, \sigma)\) is usually called “the write-back phase” and serves two purposes here:

- It is used instead of the “confirming” phase to prevent the “I still work here” and the “slow-reader” attacks. Indeed, if the configuration was superseded during the execution of Get\((a, \sigma)\), Set\((a, \sigma)\) will not succeed because it will not be able to gather a quorum of signed replies in the same configuration;
- Also, it is used to order the calls to Read\((a, \sigma)\) and to guarantee the MR-Atomicity property. Intuitively, if some correct process successfully completed Set\((a, \sigma)\) strictly before some other correct process invoked Get\((a, \sigma)\), the later process will receive a value that is not smaller than \( v \) (unless the “slow reader” attack happens).

Replica implementation. The replica implementation (Algorithm 9) essentially follows the DBLA guidelines (Algorithms 4 and 5), except that the replica handles client requests specific to Max-Register (lines 196–206). The only other difference is that
Algorithm 8 Dynamic Max-Register: code for client $p$

**Parameters:**

159. Lattice of configurations $C$ and the initial configuration $C_{init} \in C$
160. Set of values $V$ and the initial value $V_{init} \in V$
161. Boolean functions $\text{VerifyHistory}(h, \sigma)$ and $\text{VerifyInputValue}(v, \sigma)$

**Global variables:**

162. $\text{history} \subseteq C$, initially $\{C_{init}\}$
163. $\text{seqNum} \in \mathbb{Z}$, initially 0

**Auxiliary functions:** $\text{HighestConf}(h)$, $\text{FSVerify}$ (see Section 2)

164. **operation Read()**
165. repeat
166.   let $(\text{readOk}, \{v, \sigma\}) = \text{Get}()$
167.   let $\text{success} = \text{if readOk then Set}(v, \sigma) \text{ else false}$
168. until $\text{success}$
169. return $\text{seqNum} \forall$

170. **operation Write$(v, \sigma)$**
171. repeat let $\text{success} = \text{Set}(v, \sigma)$
172. until $\text{success}$

173. **operation UpdateHistory$(h, \sigma)$**
174. $\text{RB-Broadcast (NewHistory, } h, \sigma)$

175. **procedure Set$(v, \sigma)$**
176. $\text{seqNum} \leftarrow \text{seqNum} + 1$
177. let $C = \text{HighestConf}(\text{history})$
178. send $(\text{Set}, v, \text{seqNum}, C)$ to replicas($C$)
179. wait for $(\text{HighestConf}(\text{history}) \neq C) \lor (\text{replies from } Q \in \text{quorums}(C) \text{ with valid signatures})$
180. return $\text{HighestConf}(\text{history}) \neq C$

181. **procedure Get()**
182. $\text{seqNum} \leftarrow \text{seqNum} + 1$
183. let $C = \text{HighestConf}(\text{history})$
184. send $(\text{Get}, \text{seqNum}, C)$ to replicas($C$)
185. wait for $(\text{HighestConf}(\text{history}) \neq C) \lor (\text{replies from } Q \in \text{quorums}(C))$
186. if $\text{HighestConf}(\text{history}) \neq C$ then return $(\text{false}, \bot)$
187. else return $(\text{true}, \text{maximal verifiable input value among received})$

188. **upon RB-deliver (NewHistory, } h, \sigma) \text{ from any sender**}
189. if $\text{VerifyHistory}(h, \sigma) \land \text{history} \subseteq h$ then $\text{history} \leftarrow h$

in handling the $\text{UpdateRead}$ and $\text{UpdateReadResp}$ messages (lines 209–214), the replicas exchange $\sigma_{curr}$ and $\sigma_{curr}$ instead of $\text{curVals}$, as in DBLA.

B.2 Proof of correctness

Since our Dynamic Max-Register implementation uses the same state transfer protocol as DBLA, most proofs from Section A that apply to DBLA, also apply to DMR (with some minor adaptations). Here we provide only the statements of such theorems, without repeating the proofs. Then we introduce several theorems specific to DMR and sketch the proofs.

**Safety:**

**Lemma B.1 (Candidate configurations).**

(1) Each candidate configuration is present in some verifiable history.

(2) There is a finite number of candidate configurations.

(3) All candidate configurations are comparable with "$\subseteq\$".

**Lemma B.2 (Tentative configurations).**

(1) No correct client will ever make a request to a tentative configuration.

(2) Tentative configurations cannot be installed.

(3) A correct process will never invoke $\text{FSVerify}$ with timestamp $\text{height}(C)$ for any tentative configuration $C$.

(4) A correct replica will never broadcast any message via the uniform reliable broadcast primitive in a tentative configuration.

**Lemma B.3.** If $C \subseteq \text{HighestConf}(h)$, where $C$ is a pivotal configuration and $h$ is the local history of a correct process, then $C \in h$.

**Theorem B.4 (Dynamic Validity).** Our implementation of DMR satisfies Dynamic Validity. I.e., only a candidate configuration can be installed.
Algorithm 9 Dynamic Max-Register: code for replica \( r \)

Parameters: same as in Algorithm 8.

Global variables:

\[\begin{align*}
\text{history} & \subseteq C, \text{initially \{Cinit\}} \\
\nu_{\text{curr}} & \in \mathbb{V}, \text{initially Vinit} \\
\sigma_{\text{curr}} & \in \Sigma, \text{initially } \sigma_{\text{init}} \\
C_{\text{curr}} & \in C, \text{initially Cinit} \\
C_{\text{init}} & \in C, \text{initially Cinit} \\
inStateTransfer & \in \{true, false\}, \text{initially false}
\end{align*}\]

Auxiliary functions: \( \text{HighestConf}(h) \), \( \text{FSSign}(m, t) \), \( \text{UpdateFSKey}(t) \) (see Section 2)

upon receive \( \langle \text{Get}, sn, C \rangle \) from client \( c \)

190: \quad \text{wait for } C = C_{\text{init}} \lor \text{HighestConf(history)} \notin C
191: \quad \text{if } C = \text{HighestConf(history)} \text{ then}
192: \quad \quad \text{send } \langle \text{GetResp}, \nu_{\text{curr}}, \sigma_{\text{curr}}, sn \rangle \text{ to } c
193: \quad \quad \text{else ignore the message}

upon receive \( \langle \text{Set}, v, \sigma, sn, C \rangle \) from client \( c \)

194: \quad \text{wait for } C = C_{\text{init}} \lor \text{HighestConf(history)} \notin C
195: \quad \text{if } C = \text{HighestConf(history)} \land \text{VerifyInputValue}(v, \sigma) \text{ then}
196: \quad \quad \text{if } v > \nu_{\text{curr}} \text{ then } \langle \nu_{\text{curr}}, \sigma_{\text{curr}} \rangle \leftarrow \langle v, \sigma \rangle
197: \quad \quad \text{send } \langle \text{SetResp}, \text{FSSign}(\langle c, sn \rangle, \text{height}(C)), sn \rangle \text{ to } c
198: \quad \quad \text{else ignore the message}

\[\text{\( \triangleright \) State transfer}\]

199: \quad \text{upon } C_{\text{curr}} \neq \text{HighestConf}(\{C \in \text{history} | r \in \text{replicas}(C)\}) \land \text{inStateTransfer}
200: \quad \text{Same as for DBLA (lines 85–96)}

upon receive \( \langle \text{UpdateRead}, C, sn \rangle \) from replica \( r' \)

201: \quad \text{wait for } C \in \text{HighestConf(history)}
202: \quad \text{send } \langle \text{UpdateReadResp}, \nu_{\text{curr}}, \sigma_{\text{curr}}, sn \rangle \text{ to } r'

upon receive \( \langle \text{UpdateReadResp}, v, \sigma, sn \rangle \) from replica \( r' \)

203: \quad \text{if } \text{VerifyInputValue}(v, \sigma) \land v > \nu_{\text{curr}} \text{ then}
204: \quad \quad \langle \nu_{\text{curr}}, \sigma_{\text{curr}} \rangle \leftarrow \langle v, \sigma \rangle

205: \quad \text{upon } \text{RB-deliver}(\langle \text{NewHistory}, h, \sigma \rangle) \text{ from any sender}
206: \quad \text{Same as for DBLA (lines 97–100)}

upon URB-deliver(\langle \text{UpdateComplete} \rangle) \text{ in } C \text{ from quorum } Q \in \text{quorums}(C)
207: \quad \text{Same as for DBLA (lines 106–112)}

Lemma B.5 (Key update). If a pivotal configuration \( C \) is superseded, then there is no quorum of replicas in that configuration capable of signing messages with timestamp \( \text{height}(C) \), i.e., \( \exists Q \in \text{quorums}(C) \) s.t. \( \forall r \in Q : \text{st}_r \leq \text{height}(C) \).

We say that a correct client completes its operation in configuration \( C \) iff at the moment when the client completes its operation, the highest configuration in its local history is \( C \).

Lemma B.6 (State transfer correctness).

If some correct process completed \( \text{Write}(v, \sigma) \) in \( C \) or received \( v \) from \( \text{Read()} \) operation completed in \( C \), then for each active installed configuration \( D \) such that \( C \sqsubset D \), there is a quorum \( Q_D \in \text{quorums}(D) \) such that for each correct replica in \( Q_D \), \( \nu_{\text{curr}} \geq v \).

The following lemma is the first lemma specific to DMR.

Lemma B.7 (MR-Atomicity in one configuration).

If some correct process \( p \) completed \( \text{Write}(v, \sigma) \) in \( C \) or received \( v \) from \( \text{Read()} \) operation completed in \( C \) strictly before some correct process \( q \) invoked \( \text{Read()} \) and \( q \) completed its operation in \( C \), then the value returned to \( q \) is greater than or equal to \( v \).

Proof. Recall that \( \text{Read()} \) operation consists of repeated calls to two procedures: \( \text{Get()} \) and \( \text{Set}(\ldots) \). If process \( q \) successfully completed \( \text{Set}(\ldots) \) in configuration \( C \), then, by the use of forward-secure signatures, configuration \( C \) was active during the execution of \( \text{Get()} \) that preceded the call to \( \text{Set} \). This also means that configuration \( C \) was active during the execution of \( \text{Set}(v, \sigma) \) by process \( p \), since it was before process \( q \) started executing its request. By the quorum intersection property, process \( q \) must have received \( v \) or a greater value from at least one correct replica. \( \square \)
Theorem B.8 (MR-Atomicity). Our implementation of DMR satisfies the MR-Atomicity property. If some correct process \( p \) completed \( \text{Write}(v, \sigma) \) or received \( v \) from \( \text{Read()} \) strictly before some correct process \( q \) invoked \( \text{Read()} \), then the value returned to \( q \) must be greater than or equal to \( v \).

Proof. Let \( C \) (resp., \( D \)) be the highest configuration in \( p \)'s (resp., \( q \)'s) local history when it completed its request. Also, let \( v \) (resp., \( u \)) be the value that \( p \) (resp., \( q \)) passed to the last call to \( \text{Set}(\ldots) \) (note that both \( \text{Read()} \) and \( \text{Write}(\ldots) \) call \( \text{Set}(\ldots) \)).

If \( C = D \), then \( u \geq v \) by Lemma B.7.

Suppose, for contradiction, that \( D \subset C \). Since correct replicas do not reply to user requests in a configuration until this configuration is installed (line 197), configuration \( C \) had to be installed before \( p \) completed its request. By Lemma B.5 (Key Update), this would mean that \( q \) would not be able to complete \( \text{Set}(\ldots) \) in \( D \)—a contradiction.

The remaining case is when \( C \subset D \). In this case, by Lemma B.6, the quorum intersection property, and the use of forward-secure signatures in \( \text{Set}(\ldots) \), \( q \) received \( v \) or a greater value from at least one correct replica during the execution of \( \text{Get}(\ldots) \). Therefore, in this case \( u \) is also greater than or equal to \( v \).

\[ \square \]

Theorem B.9 (DMR safety). Our implementation satisfies the safety properties of DMR: MR-Validity, MR-Atomicity, and Dynamic Validity.

Proof. MR-Validity follows directly from the implementation: correct clients only return verifiable input values from \( \text{Get}(\ldots) \) (line 187). MR-Atomicity follows directly from Theorem B.8. Dynamic Validity follows from Theorem B.4.

\[ \square \]