Reconfigurable Atomic Transaction Commit

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ABSTRACT

Modern data stores achieve scalability by partitioning data into shards and fault-tolerance by replicating each shard across several servers. A key component of such systems is a Transaction Certification Service (TCS), which atomically commits a transaction spanning multiple shards. Existing TCS protocols require $2f + 1$ crash-stop replicas per shard to tolerate f failures. In this paper we present atomic commit protocols that require only $f + 1$ replicas and reconfigure the system upon failures using an external reconfiguration service. We furthermore rigorously prove that these protocols correctly implement a recently proposed TCS specification. We present protocols in two different models—the standard asynchronous message-passing model and a model with Remote Direct Memory Access (RDMA), which allows a machine to access the memory of another machine over the network without involving the latter's CPU. Our protocols are inspired by a recent FARM system for RDMA-based transaction processing. Our work codifies the core ideas of FARM as distributed TCS protocols, rigorously proves them correct and highlights the trade-offs required by the use of RDMA.

CCS CONCEPTS

• **Theory of computation → Distributed algorithms**; *Distributed computing models*;

KEYWORDS

Atomic commit, vertical Paxos, RDMA.

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1 INTRODUCTION

Modern data stores are often required to manage massive amounts of data while providing stringent transactional guarantees to their users. They achieve scalability by partitioning data into independently managed *shards* (aka *partitions*) and fault-tolerance by replicating each shard across a set of servers [7, 11, 23]. Such data stores often use optimistic concurrency control [27], where a transaction is first executed speculatively, and the results (e.g., read and write sets) are then *certified* to determine whether the transaction can commit or must abort because of a conflict with concurrent transactions. The certification is implemented by a *Transaction Certification Service (TCS)*, which accepts a stream of transactions and outputs decisions based on a given *certification function*, defining the concurrency-control check for the desired isolation level. TCS is the most challenging part of transaction processing in systems with the above architecture, since it requires solving a distributed agreement problem among the replicated shards participating in the transaction. This agreement problem has been recently formalized as the *multi-shot commit problem* [6], generalizing the classical *atomic commit problem* [10] to more faithfully reflect the requirements of modern transaction processing systems (we review the new problem statement in §2).

Most existing solutions to the TCS problem require replicating each shard among $2f + 1$ replicas to tolerate f crash-stop failures within each shard [7, 11, 16, 28], which allows using a replication protocol such as Paxos [17]. This is expensive: if transaction data are written to all replicas of the shard, only $f + 1$ replicas are needed for the data to survive failures. Since, in this case even a single replica failure will block transaction processing, to recover we need to *reconfigure* the system, i.e., change its membership to replace failed replicas with fresh ones. Unfortunately, processes concurrently deciding to reconfigure the system need to be able to agree on the next configuration; this reduces to solving consensus, which again requires $2f + 1$ replicas [18]. The way out of this conundrum is to use a separate *configuration service* with $2f + 1$ replicas to perform consensus on the configuration. In this way, we use $2f + 1$ replicas only to store the small amount of information about the configuration and $f + 1$ replicas to store the actual data. This *vertical approach* [19], which layers replication on top of a configuration service, has been used by a number of practical systems [1, 5, 9]. It is particularly suitable for deployment in local-area networks, where the configuration service can be reached quickly.

In this paper we propose the first rigorously proven protocols for implementing a TCS in a vertical system, with $f + 1$ replicas per shard and an external configuration service. We present protocols in two different models—the standard asynchronous message-passing model (§3) and a model with Remote Direct Memory Access (RDMA), which allows a machine to access the memory of another machine over the network without involving the latter's CPU (§5). Our protocols are parametric in the isolation level provided, and we prove that they correctly implement the TCS specification from the multi-shot commit problem [6] (§4).

Our work complements and takes its inspiration from a recent FARM system [9, 24]—a transaction processing system that achieves impressive scalability and availability by exploiting RDMA and the vertical approach. FARM currently forms the core of a graph database used to serve some of search queries in Microsoft Bing. It is a complex system that includes a number of optimizations, both specific to RDMA and not. FARM's design was presented without a rigorous proof of correctness, and it did not highlight which features

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are motivated by the use of RDMA and which are inherent to the vertical approach. Our work provides a theoretical complement to FARM: we codify its core ideas as distributed transaction commit protocols and rigorously prove them correct with respect to the TCS specification. By basing our protocols on a principled footing, we are also able to provide better fault-tolerance guarantees than FARM. Finally, by presenting two related protocols using message passing and RDMA, we determine the trade-offs required by the use of RDMA.

In more detail, a straightforward way to implement TCS is using the classical *two-phase commit (2PC)* protocol [12]. Since 2PC is not fault-tolerant, we can make each shard simulate a reliable process in 2PC using a replication protocol such as Paxos [7, 11]. This vanilla approach requires every 2PC action to be replicated using Paxos, which results in a high latency (7 message delays to learn a decision on a transaction [22]) and a high load on Paxos leaders. To improve on this, our protocol combines 2PC and Vertical Paxos [19] into one coherent protocol, thereby minimizing the latency and load on Paxos leaders. Upon a failure inside a shard, we use the reconfiguration service to replace the failed replicas, as in Vertical Paxos. This reconfiguration interacts nontrivially with the 2PC part of the protocol: e.g., reconfiguration may lead to losing undecided transactions that affected 2PC computations of decisions on other transactions—a behavior that we nevertheless show to be correct. Finally, we show that the price of exploiting RDMA to efficiently write transaction data to replicas is that reconfiguration has to be performed globally, instead of per-shard: when reconfiguring a shard, we have to ensure that the whole system is aware of the configuration before activating it.

2 TRANSACTION CERTIFICATION SERVICE

Service interface and certification functions. A *Transaction Certification Service (TCS)* is meant to be used in the context of transactional processing systems with optimistic concurrency control [27], where transactions are first executed speculatively, and the results are submitted for certification to the TCS. We start by reviewing its specification proposed in [6]. Clients invoke the TCS using requests of the form $\mathtt{certify}(t, l)$, where $t \in \mathcal{T}$ is a unique transaction identifier and $l \in \mathcal{L}$ is the transaction payload, which carries the results of the optimistic execution of the transaction (e.g., read and write sets). Responses of the service are of the form $\mathtt{decide}(t, d)$, where $d \in \mathcal{D} = \{\textsc{abort}, \textsc{commit}\}$. A TCS is specified using a *certification function* $f : 2^{\mathcal{L}} \times \mathcal{L} \to \mathcal{D}$, which encapsulates the concurrency-control policy for the desired isolation level. The result $f(L, l)$ is the decision for the transaction with payload l given the set of payloads L of the previously committed transactions. We require f to be *distributive* in the following sense:

$$\forall L_1, L_2, l.\, f(L_1 \cup L_2, l) = f(L_1, l) \sqcap f(L_2, l), \tag{1}$$

where $\sqcap$ is such that $\textsc{commit} \sqcap \textsc{commit} = \textsc{commit}$ and $d \sqcap \textsc{abort} = \textsc{abort}$ for any d. This requirement is justified by the fact that common definitions of $f(L, l)$ check l for conflicts against each transaction in L separately.

As an example, consider a transactional system managing *objects* from Obj with values from Val, where transactions can execute reads and writes on the objects. The objects are associated with a totally ordered set Ver of *versions*. Then the payload of a transaction t is a triple $\langle R, W, V_c \rangle$. Here the *read set* $R \subseteq \mathrm{Obj} \times \mathrm{Ver}$ is the set of objects with their versions that t read, which contains one version per object. The *write set* $W \subseteq \mathrm{Obj} \times \mathrm{Val}$ is the set of objects with their values that t wrote, which contains one value per object. We require that any object written has also been read: $\forall (x, \_) \in W.\, (x, \_) \in R$. Finally, the *commit version* $V_c \in \mathrm{Ver}$ is the version to be assigned to the writes of t. We require this version to be higher than any of the versions read: $\forall (\_, v) \in R.\, V_c > v$. Given this domain of transactions, the following certification function encapsulates the classical concurrency-control policy for serializability [27]: $f(L, \langle R, W, V_c \rangle) = \textsc{commit}$ iff none of the versions in R have been overwritten by a transaction in L, i.e.,

$$\forall x, v.\, (x, v) \in R \implies (\forall (\_, W', V'_c) \in L.\, (x, \_) \in W' \implies V'_c \le v). \tag{2}$$

TCS specification. We represent TCS executions using *histories*—sequences of $\mathtt{certify}$ and $\mathtt{decide}$ actions such that every transaction appears at most once in $\mathtt{certify}$, and each $\mathtt{decide}$ is a response to exactly one preceding $\mathtt{certify}$. For a history h we let $\mathrm{act}(h)$ be the set of actions in h. For actions $a, a' \in \mathrm{act}(h)$, we write $a \prec_h a'$ when a occurs before a' in h. A history h is *complete* if every $\mathtt{certify}$ action in it has a matching $\mathtt{decide}$ action. A complete history is *sequential* if it consists of pairs of $\mathtt{certify}$ and matching $\mathtt{decide}$ actions. A transaction t *commits* in a history h if h contains $\mathtt{decide}(t, \textsc{commit})$. We denote by $\mathrm{committed}(h)$ the projection of h to actions corresponding to the transactions that are committed in h. For a complete history h, a *linearization* ℓ of h [14] is a sequential history such that h and ℓ contain the same actions and

$$\forall t, t'.\, \mathtt{decide}(t, \_) \prec_h \mathtt{certify}(t', \_) \implies \mathtt{decide}(t, \_) \prec_\ell \mathtt{certify}(t', \_).$$

A complete sequential history h is *legal* with respect to a certification function f, if its decisions are computed according to f:

$$\forall t, l, d.\, \mathtt{certify}(t, l), \mathtt{decide}(t, d) \in \mathrm{act}(h) \implies d = f(\{l' \mid \mathtt{certify}(t', l') \in \mathrm{act}(h) \land \mathtt{decide}(t', \textsc{commit}) \prec_h \mathtt{decide}(t, d)\}, l).$$

A history h is *correct* with respect to f if $h \mid \mathrm{committed}(h)$ has a legal linearization. A TCS implementation is *correct* with respect to f if so are all its histories.

A TCS implementation satisfying the above specification can be readily used in a transaction processing system. For example, consider the domain of transactions defined earlier. A typical system based on optimistic concurrency control will ensure that transactions submitted for certification only read versions written by previously committed transactions. A history produced by such a system that is correct with respect to certification function (2) is also serializable [6]. Hence, a TCS correct with respect to this certification function can indeed be used to implement serializability.

Shard-local certification functions. We are interested in TCS implementations in systems where the data are partitioned into shards from a set $\mathcal{S}$. In such systems TCS is usually implemented using a variant of the classical *two-phase commit protocol (2PC)* [12].

In this protocol each shard s receiving a transaction for certification first *prepares* it, i.e., performs a local concurrency-control check and accordingly votes to commit or abort the transaction. The votes on the transaction by different shards are aggregated, and the final decision is then distributed to all shards: the transaction can commit only if all votes are commit. When a shard s votes on a transaction, it does not have information about all transactions in the system, but only those that concern it. Hence, the votes are computed using not the global certification function f, but *shard-local certification functions* [6], which check for conflicts only on objects managed by the shard and correspondingly take as parameters only the parts of the transaction payloads relevant to the shard: for a payload l we denote this by $l \mid s$. For example, let Obj_s be the set of objects managed by a shard s. For a payload $l = \langle R, W, V_c \rangle$ of the form given above, we let $l \mid s = \langle R^s, W^s, V_c \rangle$, where $R^s = \{(x, \_) \in R \mid x \in \mathrm{Obj}_s\}$ and $W^s = \{(x, \_) \in W \mid x \in \mathrm{Obj}_s\}$. There are two shard-local functions, $f_s : 2^{\mathcal{L}} \times \mathcal{L} \to \mathcal{D}$ and $g_s : 2^{\mathcal{L}} \times \mathcal{L} \to \mathcal{D}$. As its first argument f_s takes the set of shard-relevant payloads of transactions that previously committed at the shard, and g_s the set of such payloads for transactions that have been prepared to commit. As their second argument, the functions take the part of the payload of the transaction being certified relevant to the shard. We require that these functions are distributive, similarly to (1).

For example, the shard-local certification functions for serializability are defined as follows: $f_s(L, \langle R, W, V_c \rangle) = \textsc{commit}$ iff

$$\forall x \in \mathrm{Obj}_s.\, \forall v.\, (x, v) \in R \implies (\forall \langle \_, W', V'_c \rangle \in L.\, (x, \_) \in W' \implies V'_c \leq v),$$

and $g_s(L, \langle R, W, V_c \rangle) = \textsc{commit}$ iff

$$\forall x \in \mathrm{Obj}_s.\, \forall v.\, ((x, \_) \in R \implies (\forall \langle \_, W', \_ \rangle \in L.\, (x, \_) \notin W')) \wedge ((x, \_) \in W \implies (\forall \langle R', \_, \_ \rangle \in L.\, (x, \_) \notin R')).$$

The function f_s certifies a transaction t against previously committed transactions similarly to the certification function (2) for serializability, but taking into account only the objects managed by the shard s. The function g_s certifies t against transactions prepared to commit, and its check is stricter than that of f_s. In our example, the function g_s aborts a transaction t if: *(i)* it read an object written by a transaction t' prepared to commit; or *(ii)* it writes to an object read by a transaction t' prepared to commit. This reflects the behaviour of typical implementations, which upon preparing a transaction acquire read locks on its read set and write locks on its write set, and abort the transaction if the locks cannot be acquired.

For a sharded TCS implementation to be correct, shard-local functions have to *match* the global certification function, i.e., perform similar conflict checks. We formalize the required conditions as follows. Assume a function shards : $\mathcal{T} \to 2^{\mathcal{S}}$ that determines the shards that need to certify a transaction with a given identifier, which are usually the shards storing the data the transaction accesses. We also assume a distinguished *empty* payload $\varepsilon \in \mathcal{L}$ such that $\forall s, L.\, f_s(L, \varepsilon) = \textsc{commit}$. For example, for a payload $l = \langle R, W, \_ \rangle$ of the form given above, $l = \varepsilon$ is such that $R = \emptyset$ and $W = \emptyset$. We require that for a transaction $t \in \mathcal{T}$ with payload $l \in \mathcal{L}$, for each shard $s \notin \mathrm{shards}(t)$, we have $l \mid s = \varepsilon$. We further lift the $\mid$ operator to sets of payloads: for any $L \subseteq \mathcal{L}$ we let $(L \mid s) = \{(l \mid s) \mid l \in L\}$. Then we require that global and local certification functions match as follows:

$$\forall l \in \mathcal{L}.\, \forall L \subseteq \mathcal{L}.\, f(L, l) = \textsc{commit} \iff \forall s \in \mathcal{S}.\, f_s((L \mid s), (l \mid s)) = \textsc{commit}. \quad (3)$$

Finally, for each shard $s \in \mathcal{S}$, the two functions f_s and g_s are required to be related to each other as follows [6]:

$$\forall l \in \mathcal{L}.\, L \subseteq \mathcal{L}.\, g_s(L, l) = \textsc{commit} \implies f_s(L, l) = \textsc{commit}; \quad (4)$$

$$\forall l, l' \in \mathcal{L}.\, g_s(\{l\}, l') = \textsc{commit} \implies f_s(\{l'\}, l) = \textsc{commit}. \quad (5)$$

Property (4) requires the conflict check performed by g_s to be no weaker than the one performed by f_s. Property (5) requires a form of commutativity: if a transaction with payload l' is allowed to commit after a still-pending transaction with payload l, then the latter would be allowed to commit after the former.

3 ATOMIC COMMIT PROTOCOL

System model. We consider an asynchronous message-passing system consisting of a set of processes $\mathcal{P}$ which may fail by crashing, i.e., permanently stopping execution. We assume that processes are connected by reliable FIFO channels: messages are delivered in FIFO order, and messages between non-faulty processes are guaranteed to be eventually delivered. A function client : $\mathcal{T} \to \mathcal{P}$ determines the client process that issued a given transaction.

Each shard $s \in \mathcal{S}$ is managed by a group of *replica* processes, whose membership can change over time. For simplicity, we assume that the groups of replica processes managing different shards are disjoint. Each shard moves through a sequence of *configurations*, determining its membership. *Reconfiguration* is the process of changing the configuration of a shard. In our protocols reconfiguration is initiated by a replica when it suspects another replica of failing: for simplicity we do not expose it in the TCS interface. Every member of a shard in a given configuration is either the *leader* of the shard or a *follower*. A configuration of a shard s is then a tuple $\langle e, M, p_l \rangle$ where e is the *epoch* identifying the configuration, $M \in 2^{\mathcal{P}}$ is the set of processes that manage s at e, and $p_l \in M$ is the leader of s at e.

Configurations are stored in an external *configuration service (CS)*, which for simplicity we assume to be a reliable process. In practice, this service may be implemented using Paxos-like replication over $2f + 1$ processes out of which at most f can fail (as done in systems such as Zookeeper [15]). The configuration service stores the configurations of all shards and provides three operations. An operation `compare_and_swap`$(s, e, \langle e', M, p_l \rangle)$ succeeds if the epoch of the last stored configuration of s is e; in this case it stores the provided configuration with a higher epoch $e' > e$. Operations `get_last`(s) and `get`(s, e) respectively return the last configuration of s and the configuration of s associated with a given epoch e.

Protocol preliminaries. We give the pseudocode of our protocol in Figure 1, illustrate its message flow in Figure 2 and summarize the key invariants used in its proof of correctness in Figure 3. The protocol weaves together the two-phase commit protocol across shards [12] and a Vertical Paxos-based reconfiguration protocol within each shard [19]. At any given time, a process participates in a single configuration of the shard it belongs to. The process stores the information about this configuration as well as those of other shards in several arrays: configuration epochs are stored in an array epoch $\in \mathcal{S} \rightarrow \mathbb{N}$, the current members in members $\in \mathcal{S} \rightarrow 2^{\mathcal{P}}$, and the current leader in leader $\in \mathcal{S} \rightarrow \mathcal{P}$. The entries for the shard the process belongs to give the configuration the process is in; the other entries maintain information about the configurations of the other shards. A status variable at a process records whether it is a LEADER, a FOLLOWER or is in a special RECONFIGURING state used during reconfiguration. Each process keeps track of the status of transactions in an array phase, whose entries initially store START. The transaction status changes to PREPARED when the shard determines its vote and to DECIDED when a final decision on the transaction is reached.

Failure-free case. A client submits a transaction for certification by calling the `certify` function at any replica process, which will serve as the *coordinator* of the transaction (line 1). The function takes as arguments the transaction's identifier and its payload. The transaction coordinator first sends a `PREPARE` message to the leaders of the relevant shards, which includes the payload part for each shard (line 3). The leader of a shard arranges all transactions it receives into a total *certification order*, which the leader stores in an array txn $\in \mathbb{N} \rightarrow \mathcal{T}$; a next $\in \mathbb{Z}$ variable points to the last filled slot in the array. When the leader receives a `PREPARE` message for a transaction for the first time (line 8), it appends the transaction to the certification order, stores the transaction's payload in an array payload $\in \mathbb{N} \rightarrow \mathcal{L}$, and sets the transaction's phase to PREPARED. It then computes a vote on the transaction and stores it in an array vote $\in \mathbb{N} \rightarrow \{\text{COMMIT}, \text{ABORT}\}$ (line 12). The vote is computed using the shard-local certification functions f_{s_0} and g_{s_0} to check for conflicts against transactions that have been previously committed or prepared to commit; the results are combined using the $\sqcap$ operator, so that the transaction can commit only if both functions say so. We defer the description of the cases when the leader has previously received the transaction in the `PREPARE` message (line 6) and when the payload in the message is an undefined value $\perp$ (line 14).

Our protocol next replicates the leader's decision and the transaction payload at the followers. Instead of having the leader to do this directly, the protocol delegates this task to the coordinator of the transaction. This design is used by practical systems, such as Corfu [1] and FARM [9], since it minimizes the load on the leaders, which are the main potential performance bottleneck. Instead, the network-intensive task of persisting transactions at multiple followers is spread among a number of different transaction coordinators. As we explain in the following, this optimization interacts in a nontrivial way with transaction certification. In more detail, after preparing a transaction the leader sends a `PREPARE_ACK` message to the coordinator of the transaction, which carries the leader's epoch, the transaction identifier, its position in the certification order, the payload, and the vote (line 20). Upon receiving the `PREPARE_ACK` message (line 18), the coordinator forwards the data from the `PREPARE_ACK` message to the followers in an `ACCEPT` message.

A process handles an `ACCEPT` message only if it participates in the corresponding epoch (line 22). The process stores the transaction identifier, its payload and vote, and advances the transaction's phase to PREPARED. It then sends an `ACCEPT_ACK` message to the coordinator of the transaction, confirming that the process has accepted the transaction and the vote. The certification order at a follower is always a prefix with zero or more holes of the certification order at the leader of the epoch the follower is in, as formalized by Invariant 1 (Figure 3). The holes in the prefix arise from the lack of FIFO ordering in the communication between the leader of a given epoch and its followers, as the `ACCEPT` message for a given transaction is sent to the followers by the coordinator of the transaction and not directly by the leader.

The coordinator of a transaction t acts once it receives `ACCEPT_ACK` messages for t from every follower of its shards $s \in \text{shards}(t)$ (line 26); it determines this using the configuration information it stores for every shard. The coordinator computes the final decision on t using the $\sqcap$ operator on the votes of each involved shard: the transaction can commit if all votes are commit. The coordinator then sends the final decision in `DECISION` messages to the client and to each of the relevant shards. When a process receives a decision for a transaction (line 30), it stores the decision and advances the transaction's phase to DECIDED. In a realistic implementation, at this point the process would also upcall into the transaction processing system running at its server, to inform it about the decision and allow it to apply the transaction's writes to the database if the decision is to commit.

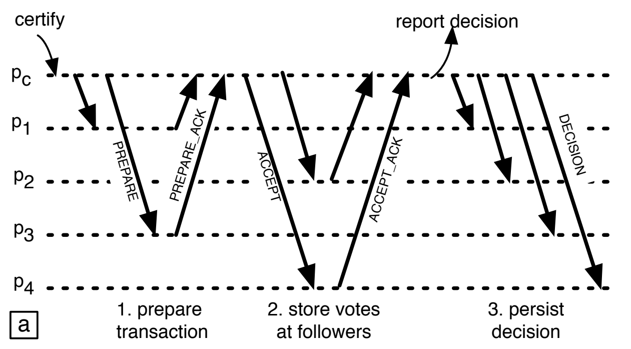
In the absence of failures, our protocol allows the client to learn a decision on a transaction in 5 message delays, instead of 7 required by vanilla protocols that use Paxos as a black-box [7, 11]. We can further reduce this to 4 by co-locating the client with the transaction coordinator. The protocol also minimizes the load on Paxos leaders, which are the main potential bottleneck: each involved leader only has to receive one `PREPARE` and one `DECISION` message, and send one `PREPARE_ACK` message.

Reconfiguration. When a failure is suspected in a shard s, any process can initiate a reconfiguration of the shard to replace failed replicas. Reconfiguration is done only in the affected shard, without disrupting others. It aims to preserve Invariant 2, which is key in proving the correctness of the protocol. This assumes that all followers in s at an epoch e have received `ACCEPT`(e, k, t, l, d) and responded to it with `ACCEPT_ACK`; in this case we say that the transaction t has been *accepted* at shard s. The invariant guarantees that the accepted transaction t will persist in epochs higher than e; this is used to prove that the protocol computes a unique decision on each transaction. The invariant also guarantees that the entries preceding t in the certification order in epochs higher than e may only contain the votes that the leader of s at epoch e took into account when computing the vote d on t (some of these votes may be missing due to the lack of FIFO order in the communication between the leader and its followers). This property is necessary to guarantee that the protocol computes decisions according to a single global certification order, as required by the TCS specification.

To ensure Invariant 2, a process performing reconfiguration first *probes* previous configurations to determine which processes are still alive and to find a process whose state contains all transactions previously accepted at the shard, which will serve the new leader. The new leader then transfers its state to the members of the new configuration, thereby *initializing* them. A variable initialized $\in \{\text{TRUE}, \text{FALSE}\}$ at a process records whether it has

```
1  function certify($t, l$)
2      forall $s \in$ shards($t$) do
3          send PREPARE($t, (l \mid s)$) to leader[$s$];

4  when received PREPARE($t, l$) from $p_j$
5      pre: status = LEADER;
6      if $\exists k.\ t$ = txn[$k$] then
7          send PREPARE_ACK(epoch[$s_0$], $s_0$, $k$, txn[$k$],
                   payload[$k$], vote[$k$]) to $p_j$
8      else
9          next ← next + 1;
10         (txn, phase)[next] ← ($t$, PREPARED);
11         if $l \neq \perp$ then
12             vote[next] ← $f_{s_0}(L_1, l) \sqcap g_{s_0}(L_2, l)$;
13             payload[next] ← $l$;
14         else
15             vote[next] ← ABORT;
16             payload[next] ← $\varepsilon$;
17         send PREPARE_ACK(epoch[$s_0$], $s_0$, next, $t$,
                   payload[next], vote[next]) to $p_j$;

18 when received PREPARE_ACK($e, s, k, t, l, d$)
19     pre: epoch[$s$] = $e$;
20     send ACCEPT($e, k, t, l, d$) to members[$s$] \ leader[$s$];

21 when received ACCEPT($e, k, t, l, d$) from $p_j$
22     pre: status = FOLLOWER ∧ epoch[$s_0$] = $e$;
23     if phase[$k$] = START then
24         (txn, payload, vote, phase)[next] ← ($t, l, d$, PREPARED);
25     send ACCEPT_ACK($s_0, e, k, t, d$) to $p_j$;

26 when for every $s \in$ shards($t$) received an
   ACCEPT_ACK($s$, epoch[$s$], $k_s, t, d_s$) from every
   $p_j \in$ members[$s$] \ leader[$s$]
27     send DECISION($t, \bigsqcap_{s \in \text{shards}(t)} d_s$) to client($t$);
28     forall $s \in$ shards($t$) do
29         send DECISION(epoch[$s$], $k_s$, $\bigsqcap_{s \in \text{shards}(t)} d_s$)
               to members[$s$];

30 when received DECISION($e, k, d$)
31     pre: status ∈ {LEADER, FOLLOWER} ∧ epoch[$s_0$] ≥ $e$;
32     (dec, phase)[$k$] ← ($d$, DECIDED);

33 function reconfigure($s$)
34     pre: probing = FALSE;
35     probing ← TRUE;
36     ⟨probed_epoch, probed_members, _⟩ ← get_last($s$) at CS;
37     recon_epoch ← probed_epoch + 1;
38     recon_shard ← $s$;
39     send PROBE(recon_epoch) to probed_members;

40 when received PROBE($e$) from $p_j$
41     pre: $e$ ≥ new_epoch;
42     status = RECONFIGURING;
43     new_epoch ← $e$;
44     send PROBE_ACK(initialized, $e$, $s_0$) to $p_j$;

45 when received PROBE_ACK(TRUE, recon_epoch,
                   recon_shard) from $p_j$
46     pre: probing = TRUE;
47     probing ← FALSE;
48     $M$ ← compute_membership();
49     var $r$ ← compare_and_swap(recon_shard,
           recon_epoch − 1, ⟨recon_epoch, $M$, $p_j$⟩) at CS;
50     if $r$ then send NEW_CONFIG(recon_epoch, $M$) to $p_j$;

51 non-deterministically when received
   PROBE_ACK(FALSE, recon_epoch, recon_shard) from
   $p_j$ ∈ probed_members and no
   PROBE_ACK(TRUE, recon_epoch, recon_shard)
52     pre: probing = TRUE;
53     probed_epoch ← probed_epoch − 1;
54     probed_members ← get(recon_shard,
                   probed_epoch) at CS;
55     send PROBE(recon_epoch) to probed_members;

56 when received NEW_CONFIG(new_epoch, $M$) from $p_j$
57     status = LEADER;
58     (epoch, members, leader)[$s_0$] ← ($e, M, p_i$);
59     next ← max{$k$ | phase[$k$] ≠ START};
60     send NEW_STATE(new_epoch, $M$, txn, payload, vote,
               dec, phase) to $M \setminus p_i$;

61 when received NEW_STATE($e, M, txn, payload, vote,$
               $dec, phase$) from $p_j$
62     pre: $e$ ≥ new_epoch;
63     initialized ← TRUE;
64     status = FOLLOWER;
65     (epoch, members, leader)[$s_0$] ← ($e, M, p_j$);
66     (txn, payload, vote, dec, phase) ←
           ($txn, payload, vote, dec, phase$);

67 when received CONFIG_CHANGE($s, e, M, p_l$) from CS
68     pre: epoch[$s$] < $e$ ∧ $s \neq s_0$;
69     (epoch, members, leader)[$s$] ← ($e, M, p_l$);

70 function retry($k$)
71     pre: phase[$k$] = PREPARED;
72     forall $s \in$ shards(txn[$k$]) do
73         send PREPARE(txn[$k$], $\perp$) to leader[$s$];
```

Figure 1: Atomic commit protocol at a process p_i in a shard s_0. At line 12 we let L_1 = {payload[k] | k < next ∧ phase[k] = DECIDED ∧ dec[k] = COMMIT}; L_2 = {payload[k] | k < next ∧ phase[k] = PREPARED ∧ vote[k] = COMMIT}.



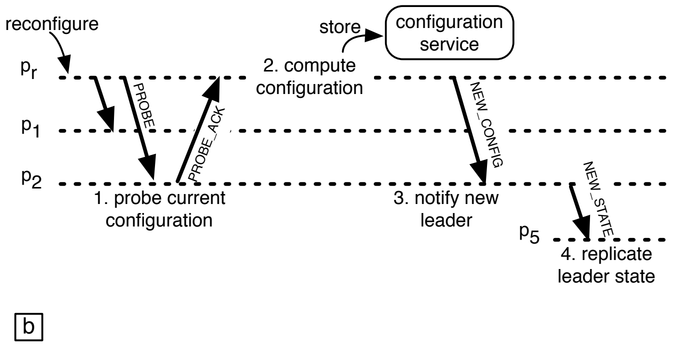


Figure 2: Illustrations of the behavior of the atomic commit protocol during (a) failure-free operation; (b) reconfiguration. In both cases p_1 is the initial leader of shard s_1 and p_2 its follower; p_3 is the leader of shard s_2 and p_4 its follower. In (b), after the reconfiguration of s_1, process p_2 becomes the leader of s_1 and a fresh process p_5 becomes its follower.

1. If a process p_i receives $\mathtt{ACCEPT}(e, k, t, l, d)$ and replies with ACCEPT_ACK, then after this and while $\mathsf{epoch}[s] = e$ at p_i, we have $\mathsf{txn}\rfloor_k \prec \mathit{txn}\rfloor_k$, $\mathsf{vote}\rfloor_k \prec \mathit{vote}\rfloor_k$ and $\mathsf{payload}\rfloor_k \prec \mathit{payload}\rfloor_k$, where *txn*, *vote* and *payload* were the values of the arrays txn, vote and payload at the leader of s at e when it sent the corresponding message $\mathtt{PREPARE\_ACK}(e, s, k, t, l, d)$.
2. Assume that all followers in s at e received $\mathtt{ACCEPT}(e, k, t, l, d)$ and responded to it with ACCEPT_ACK, and at the time the leader of s at e sent the corresponding message $\mathtt{PREPARE\_ACK}(e, s, k, t, l, d)$ it had $\mathsf{txn}\rfloor_k = \mathit{txn}$, $\mathsf{vote}\rfloor_k = \mathit{vote}$ and $\mathsf{payload}\rfloor_k = \mathit{payload}$. Whenever at a process in s we have $\mathsf{epoch}[s] = e' > e$, we also have $\mathsf{txn}\rfloor_k \prec \mathit{txn}$, $\mathsf{vote}\rfloor_k \prec \mathit{vote}$ and $\mathsf{payload}\rfloor_k \prec \mathit{payload}$.
3. After a process receives $\mathtt{PROBE}(e)$ and replies with $\mathtt{PROBE\_ACK}(\_, e, s)$, it will never send $\mathtt{ACCEPT\_ACK}(s, e', \_, \_, \_)$ with $e' < e$.
4. (a) For any messages $\mathtt{DECISION}(\_, k, d_1)$ and $\mathtt{DECISION}(\_, k, d_2)$ sent to processes in the same shard, we have $d_1 = d_2$.
   (b) For any messages $\mathtt{DECISION}(t, d_1)$ and $\mathtt{DECISION}(t, d_2)$ sent, we have $d_1 = d_2$.
5. Assume that all followers in s at e received $\mathtt{ACCEPT}(e, k, t, l, d)$ and replied with ACCEPT_ACK. Let the process p_i be a member of s at e' such that $e' < e$. If p_i is not a member of s at e then p_i cannot be member of s at any $e'' > e$.

Figure 3: Key invariants of the atomic commit protocol. Given a sequence α, we let $\mathsf{length}(\alpha) = \max\{k \mid \alpha[k] \neq \bot\}$ and let $\alpha\rfloor_k$ be the prefix of α of length k. Given a second sequence β, we let $\beta \prec \alpha$ if $\mathsf{length}(\alpha) = \mathsf{length}(\beta) \wedge (\forall k \leq \mathsf{length}(\beta).\ \beta[k] \neq \bot \implies \beta[k] = \alpha[k])$.

ever been initialized. Our protocol guarantees that a shard can become *operational*, i.e., start accepting transactions, only after all its members have been initialized.

The probing phase is complicated by the fact that there may be a series of failed reconfiguration attempts, where the new leader fails before initializing all its followers. Hence, probing requires traversing epochs from the current one down, skipping epochs that are not operational. Probing selects as the new leader the first initialized process it encounters during this traversal; we can show that this process is guaranteed to know about all transactions accepted at the shard, and thus making it the new leader will preserve Invariant 2 (§4).

In more detail, a process p_r initiates a reconfiguration of a shard s by calling $\mathtt{reconfigure}(s)$ (line 33). The process picks an epoch number recon_epoch higher than the epoch of s stored in the configuration service and then starts the probing phase, as marked by the flag probing. The process p_r keeps track of the shard being reconfigured in recon_shard, the epoch being probed in probed_epoch and the membership of this epoch in probed_members. The process initializes these variables when it first reads the current configuration from the configuration service (line 36). It then sends a PROBE message to the members of the current configuration, asking them to join the new epoch recon_epoch. Upon receiving a $\mathtt{PROBE}(e)$ message (line 40), a process first checks that the proposed epoch is equal or higher than the highest epoch it has ever been asked to join, which the process stores in new_epoch (we always have $\mathsf{epoch}[s] \leq \mathsf{new\_epoch}$ at a process in s). In this case, the process sets new_epoch to e and changes its status to RECONFIGURING, which causes it to stop transaction processing. It then replies to p_r with a PROBE_ACK message, which indicates whether it has been previously initialized or not. If p_r finds a process that has previously been initialized, and hence can serve as the new leader, p_r ends probing (line 45). If p_r does not find such a process in the epoch probed_epoch and receives at least one reply PROBE_ACK from a process that has not been initialized (line 51), p_r can conclude that the epoch probed_epoch is not operational and will never become such, because it has convinced at least one of its members to join the new epoch; this is formalized by Invariant 3. In this case p_r starts probing the preceding epoch. Since no transactions could have been accepted at the epoch probed_epoch, picking a new leader from an earlier epoch will not lose any accepted transactions and thus will not violate Invariant 2.

Once the probing finds a new leader p_j for the shard s (line 45), the process p_r computes the membership of the new configuration using a function compute_membership (line 48). We do not prescribe a particular implementation of this function, except that the new membership must contain the new leader p_j and may only contain the processes that replied to probing or fresh processes. The latter can be added to reach the desired level of fault tolerance. Once the new configuration is computed, p_r attempts to store it in the configuration service using a compare-and-swap operation. This succeeds only if the current epoch is still the epoch from which p_r started probing, which means that no concurrent reconfiguration occurred while p_r was probing. In this case, p_r sends a NEW_CONFIG message with the new configuration to the new leader of s.

When the new leader of s receives the NEW_CONFIG message (line 56), it sets next to the length of its sequence of transactions, epoch$[s]$ to the new epoch and status to LEADER, which allows it to start processing new transactions. It then sends a NEW_STATE message to its followers, containing its state. Upon receiving this message (line 61), a process overwrites its state with the one provided, changes its status to FOLLOWER, and sets initialized to TRUE. As part of the state update, the process also updates its epoch epoch$[s_0]$ to the new one. Hence, the process will not accept transactions from the new leader until it receives the NEW_STATE message.

When a new configuration of a shard s is persisted in the configuration service, the service sends it in a CONFIG_CHANGE message to the members of shards other than s. A process updates the locally stored configuration upon receiving this message (line 67).

Coordinator recovery. If a process that accepted a transaction t does not receive the final decision on it, this may be because the coordinator of t has failed. In this case the process may decide to become a new coordinator by executing a retry function (line 70). For this, the process just sends a PREPARE$(t, \bot)$ message to the leaders of the shards of t, carrying a special undefined value $\bot$ as the payload. If a leader receiving PREPARE$(t, \bot)$ has already certified t, it re-sends the corresponding PREPARE_ACK message to the new coordinator, including the transaction payload and vote (line 6). Otherwise, if the leader does not have the payload of t, it prepares the transaction as aborted and with an empty payload ε (line 15). In either case, the new coordinator will finish processing the transaction as usual. The above case when the transaction is aborted because the leader of a shard does not know its payload may arise when the old coordinator crashed in between sending PREPARE messages to different shards. Note that if the old coordinator was suspected spuriously and will try later to submit the transaction to a shard where it was aborted, it will just get a PREPARE_ACK message with an ABORT vote.

Our protocol allows any number of processes to become coordinators of a transaction at the same time. Nevertheless, the protocol ensures that they will all reach the same decision, even in case of reconfigurations. We formalize this by Invariant 4: part (a) ensures an agreement on the decision on the k-th transaction in the certification order at a given shard; part (b) ensures a system-wide agreement on the decision on a given transaction t. The latter part establishes that the protocol computes a unique decision on each transaction. Invariant 4 is proved as a corollary of Invariant 2.

Losing undecided transactions. Recall that our protocol uses the optimization that delegates persisting transactions at followers to coordinators [1, 9]. We now highlight how this optimization interacts with transaction certification. Because of the optimization, transactions prepared by a leader of a shard s can be persisted at followers out of order. For example, t_2 may follow t_1 in the certification order at the leader, but may be persisted at followers first. If now the leader of s and the coordinator of t_1 crashes before t_1 is persisted at followers, t_1 will be lost forever, something that is allowed by Invariant 2 (due to the use of $\prec$). In this case we lose a transaction t_1 on the basis of which the vote on the transaction t_2 was computed (e.g., the payload l_1 of t_1 was in L_2 when the vote on t_2 was computed at line 12). This does not violate correctness, since the vote on t_2 makes sense also in the context excluding t_1: due to distributivity of certification functions (§2), if t_2 was allowed to commit in the presence of t_1 ($f_s(\{l_1\}, l_2) = \text{COMMIT}$), it can also commit in its absence ($f_s(\emptyset, l_2) = \text{COMMIT}$). Note that in this case a decision on t_1 could not have been exposed to the client: otherwise t_1 could not get lost due to Invariant 2. Also note that, since we assume the transaction execution component produces payloads with read-sets containing only values written by committed transactions (§2), in the above case t_2 could not have read a value written by t_1.

4 CORRECTNESS

The next theorem states the safety of our protocol, showing that it implements the TCS specification.

THEOREM 4.1. *A transaction certification service implemented using the protocol in Figure 1 is correct with respect to a certification function f matching the shard-local certification functions f_s and g_s.*

We defer the proof to [4] and only sketch the proof of the key Invariant 2. This relies on auxiliary Invariant 5, which we prove first.

Proof sketch for Invariant 5. We prove the invariant by induction on e''. Assume that the invariant holds for all $e'' < e^*$. We now show it for $e'' = e^*$. The members of s at e^* are computed at line 48 by a reconfiguring process p_r using the compute_membership function, which returns either fresh processes or processes that responded to p_r's probing. Since p_i was a member of s at $e' < e^*$, it is not fresh; then by assumptions on compute_membership p_i must have received PROBE(e^*) from p_r and replied with PROBE_ACK$(_, e^*, s)$. The process p_r starts probing at epoch $e^* - 1$ and ends it upon receiving a PROBE_ACK(TRUE, e^*, s) message. By the induction hypothesis, p_i is not a member of s at any epoch from $e^* - 1$ down to $e + 1$. Hence, if the probing stops before reaching e, then p_i will not be a member of s at e^*, as required. Assume now that the probing reaches e. By Invariant 3, each follower in s at e must have sent ACCEPT_ACK(s, e, t) before receiving PROBE(e^*). Then any member of s at e receiving PROBE(e^*) will have initialized = TRUE. Hence, if any member of s at e replies with PROBE_ACK$(\textit{initialized}, e^*, s)$, we have that *initialized* = TRUE. Since the process p_r will not move to the preceding epoch until at least one process replies with PROBE_ACK, this means that the probing can never go beyond e. Since the process p_i is not a member of e, it cannot be included as a member of s in e^*, as required. □

Proof sketch for Invariant 2. We prove the invariant by induction on e'. Assume that the invariant holds for all $e' < e''$. We now show it for $e' = e''$ by induction on the length of the protocol execution. We only consider the most interesting transition in line 56, when a process p_i becomes a leader of s at an epoch e''. We show that after this transition at p_i we have $\text{txn}\rfloor_k \prec \textit{txn}$, $\text{vote}\rfloor_k \prec \textit{vote}$ and $\text{payload}\rfloor_k \prec \textit{payload}$.

Since p_i was chosen as the leader of s at e'', this process replied with PROBE_ACK(TRUE, e'', s) to a PROBE(e''). Therefore, p_i was a member of s at an epoch $e^* < e''$ that was being probed. Probing ends when at least one process sends a PROBE_ACK(TRUE, e'', s). From Invariant 3 and the assumption that all followers in e replied with ACCEPT_ACK to ACCEPT(e, k, t, l, d), we can conclude that probing could no have gone further than e. Hence, $e \leq e^* < e''$.

Let e_0 be the value of epoch[s] at p_i right before the transition at line 56. We have $e_0 \geq e$, as otherwise p_i would not be a member of s at e and by Invariant 5 could not be picked as the leader of s at e''. It is also easy to show that $e_0 < e''$. Hence, $e \leq e_0 < e''$.

If $e < e_0$, then by the induction hypothesis, we have $\text{txn}\rfloor_k \prec \mathit{txn}$, $\text{vote}\rfloor_k \prec \mathit{vote}$ and $\text{payload}\rfloor_k \prec \mathit{payload}$ right after the transition in line 56, as required. Assume now that $e_0 = e$. If p_i was the leader of s at e, then we trivially have $\text{txn}\rfloor_k \prec \mathit{txn}$, $\text{vote}\rfloor_k \prec \mathit{vote}$ and $\text{payload}\rfloor_k \prec \mathit{payload}$ right after the transition in line 56, as required. Otherwise, by Invariant 3, p_i must have received ACCEPT(e, k, t, l, d) and responded to it with ACCEPT_ACK(s, e, t) before the transition in line 56. Then the required follows from Invariant 1. □

We next state liveness properties of our protocol (we again defer proofs to [4]). The reconfiguration procedure in the protocol will get stuck if it cannot find an initialized process, which may happen if enough processes crash, so that all shard data is lost. We now state conditions under which this cannot happen. We associate two events with each configuration e of a shard s: *introduction* and *activation*. Introduction indicates that the configuration comes into existence and is triggered when the configuration is successfully persisted in the configuration service (line 49). Activation indicates that the configuration becomes operational and is triggered when all the followers of the configuration have processed the NEW_STATE messages sent by its leader (line 61).

Once a configuration has been activated, we say that it is *active*. We define its *lifetime* as the time interval between its introduction and when a succeeding configuration becomes active. Note that not every introduced configuration necessarily becomes active, since its leader may never complete the data transfer to the followers. To ensure our protocol is live we make the following assumption, similar to the ones made by other protocols with changing membership [3, 26].

ASSUMPTION 1. *At least one member in each configuration is non-faulty throughout the lifetime of a configuration.*

The following two theorems show that, under this assumption, a single reconfiguration makes progress.

THEOREM 4.2. *If a process p_r attempts to reconfigure a shard s and no other process attempts to reconfigure s simultaneously, then if p_r is non-faulty for long enough, it will eventually introduce a new configuration.*

THEOREM 4.3. *If a configuration of a shard s is introduced by a process p_r, then it will eventually be activated, provided no process attempts to reconfigure s simultaneously, and p_r and all the members of the configuration are non-faulty for long enough.*

Finally, the following theorem shows that in the absence of failures or reconfigurations, transaction certification makes progress.

THEOREM 4.4. *Assume that the current configuration of each shard is active, all processes are aware of the current configuration of each shard, and no reconfiguration is in progress. If a transaction is submitted for certification, then it will eventually be decided, provided no reconfiguration is attempted and all the processes belonging to the current configuration of each shard are non-faulty for long enough.*

5 EXPLOITING RDMA

We now present a variant of our protocol that uses Remote Direct Memory Access (RDMA), which follows the design of the FARM system [9, 24]. By comparing this protocol with that of §3 we highlight the trade-offs required by the use of RDMA. Due to space constraints, we defer the pseudocode of our protocol to [4] and describe the required changes in the protocol of §3 only informally.

We assume the same system model as in §2, except that processes can communicate using RDMA. This allows a machine to access the memory of another machine over the network without involving the latter's CPU, thus lowering latency. Like FARM, our protocol uses RDMA to implement a primitive for point-to-point communication between processes with the following interface. The primitive allows a sender process to reliably send a message m to a receiver process p_j (**send-rdma**(m, p_j)) by remotely writing into a specific memory region of p_j. The sender then gets an acknowledgement when the message reaches the receiver's memory (**ack-rdma**(m, p_j)), sent by the receiver's network interface card (NIC) without interrupting its CPU. The receiver is notified at a later point that a new message is available (**deliver-rdma**(m, p_j)). Hence, the guarantee provided by **ack-rdma**(m, p_j) is that the receiver will eventually deliver the message m, even if the sender crashes, since the message is already in the receiver's memory. The operation **open**(p_i) grants p_i access to a region of the caller's memory, and **close**(p_i) revokes it. Once the latter operation completes, p_i cannot send any message to the caller using **send-rdma**. Finally, we assume that the communication primitive includes another operation: **flush**. This operation blocks the caller until it has delivered all messages addressed to it that have been acknowledged by its NIC through an **ack-rdma**.

To implement the above primitive, the receiver usually keeps a circular buffer in memory for each process that may send it a message [8, 21]. The operation **send-rdma**(m, p_j) issued by a process p_i appends a message to the corresponding buffer at the receiver using RDMA writes. Receivers periodically pull messages from the buffers and deliver them to the application via **deliver-rdma**. If a buffer at a process p_j gets full, the associated sender process will not be able to send a message to p_j until the latter pulls some messages.

Following FARM, we use the above RDMA-based communication primitive in our protocol to persists votes and decisions (steps 2 and 3 of Figure 2a). This requires the following changes to the protocol in Figure 1. First, ACCEPT and DECISION messages are sent using **send-rdma** instead of **send** (lines 20 and 29). Second, the followers do not send explicit ACCEPT_ACK messages to transaction coordinators (line 25); instead, the latter act once they receive an RDMA acknowledgement **ack-rdma**. This makes the checks at lines 22 and 31 redundant, as followers cannot reject ACCEPT or DECISION messages under any circumstance. The practical rationale for these changes is that persisting a transaction t at followers using

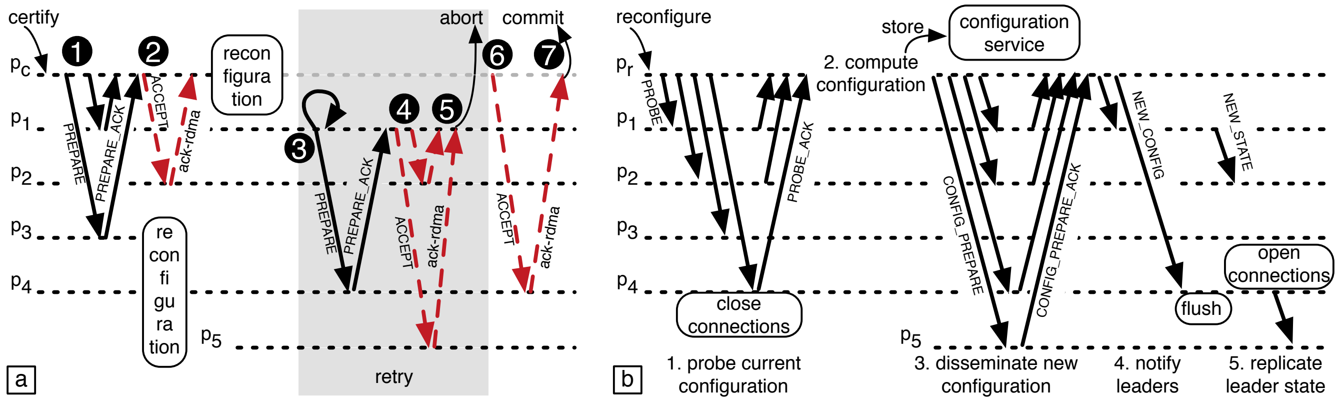


Figure 4: Illustrations of (a) a counter-example showing the need to change the reconfiguration protocol when using RDMA; (b) the changed reconfiguration protocol. Red dashed lines denote RDMA operations. In both cases p_1 is the leader of shard s_1 and p_2 its follower; p_3 is the initial leader of shard s_2 and p_4 its follower. After the reconfiguration of s_2, process p_4 is the new leader and a fresh process p_5 is its follower.

RDMA minimizes the time during which the transaction is prepared at leaders, which requires them to vote abort on all transactions conflicting with t (via the certification function g_s, §2); this results in lower abort rates [2, 9]. Transaction processing at followers (e.g., adding them to the local copy of the certification order, line 21) is done off the critical path of certification.

Unfortunately, the above changes to the failure-free path of the protocol do not preserve correctness without changes to reconfiguration, as illustrated by an example execution in Figure 4a. In this execution, two shards s_1 and s_2 are involved in the certification of a transaction t, coordinated by a process p_c from a third shard. The transaction is prepared to commit at the leaders p_1 and p_3 of both shards (step ➊), and the COMMIT vote from the leader of s_1 (p_1) is persisted at the follower p_2 using RDMA (step ➋). Before the coordinator p_c persists the vote from the leader p_3 of s_2 at the follower p_4, the leader p_3 is suspected of failure and a reconfiguration is triggered at shard s_2. This promotes the follower p_4 to a new leader and brings online a fresh follower p_5. Next, the leader p_1 of s_1 suspects the coordinator p_c of failure and triggers a reconfiguration to remove it. Once p_c is removed from its shard, p_1 retries the processing of t (step ➌, line 70 in Figure 1). The new leader p_4 of s_2 does not know about t, so this results in the transaction being aborted, because its payload at shard s_2 is thought to be lost (steps ➍ and ➎). But now the coordinator p_c, who did not actually fail and still believes s_2 is in the old configuration, finishes its processing by persisting the COMMIT vote of the old leader p_3 of s_2 at the old follower p_4, which is now the new leader of s_2 (step ➏). Since this is done via RDMA, p_4 cannot reject the vote and, thus, p_c commits the transaction (step ➐). This violates safety, as two contradictory results have been externalized. The protocol in §3 is not subject to this problem, because in that protocol the new leader p_4 of the shard s_2 would reject the ACCEPT message due to the failure of the check at line 22.

To make the RDMA-based protocol correct, we need to change the reconfiguration protocol so that *the whole system* participates in reconfiguration instead of just the affected shard. Figure 4b illustrates the message flow of the redesigned reconfiguration protocol. Processes now maintain a single epoch variable instead of a vector. The data structures maintained by the external configuration service and its interface are adjusted accordingly. Like in our previous commit protocol, the process p_r performing reconfiguration first probes previous configurations by sending PROBE messages. However, p_r now probes *all* shards. A process receiving PROBE handles it as before (line 40), but additionally closes all incoming RDMA connections using **close**, which guarantees that the process accepts no more transactions at its previous epoch. This is needed because, due to communication via RDMA, the protocol cannot longer leverage the safety check at line 22. The logic of the reconfiguring process is also changed: after this process computes the new configuration and stores it in the configuration service (line 49), the process sends a new CONFIG_PREPARE message to *all* processes in the configuration. Upon receiving CONFIG_PREPARE, a process updates its locally stored configuration and replies with a CONFIG_PREPARE_ACK message. This ensures that the whole system is aware of the new configuration before it is activated. Only after this does the reconfiguring process send a NEW_CONFIG message to the leaders of the new configuration. Upon receiving NEW_CONFIG (line 56), a leader p_l first calls **flush**. This guarantees that all the messages that have been acknowledged as having reached p_l's memory will be replicated to followers in NEW_STATE messages; this is necessary since transaction coordinators may have already externalized decisions taken based on these acknowledgements. Finally, processes open RDMA connections to all other processes in the configuration using **open**: a leader after sending NEW_STATE to its followers, and followers upon receiving NEW_STATE (line 61).

The new protocol guarantees that: (*) if a process receives an ACCEPT message for a transaction t while at epoch e, then the leader that prepared t was at epoch e when it prepared this transaction. This property is key in proving the correctness of the protocol, as it provides the same guarantees as the removed guard in line 22, which we could not leverage due to the use of RDMA. The property (*) holds because: *(i)* at any time, a process only maintain RDMA connections to the members of its current epoch; and *(ii)* before persisting a vote at a follower, the coordinator of a transaction checks that the transaction was prepared in its current epoch (line 19).

We now show how the revised reconfiguration protocol prevents the bug in Figure 4a. In this protocol, when p_c attempts to persist the COMMIT vote at p_4 (step ❻), the latter will be already aware that p_c has been removed from the system and will close the RDMA connection to it. Thus, p_c will be unable to persist the vote at p_4 (this would violate the property (*)) and will never gather enough acknowledgements to decide the transaction. Hence, no contradictory results will be externalized. We state and prove the correctness of the RDMA-based protocol in [4].

6 RELATED WORK AND DISCUSSION

Our protocols are inspired by the recent FARM system for transaction processing, which also uses $f + 1$ replicas per shard and deals with failures using reconfiguration [9, 24]. FARM was presented as a complete database system with a number of optimizations, including the use of RDMA. In contrast, our work distills the core ideas of FARM into protocols solving the well-defined transaction certification problem, parametric in the isolation level provided and rigorously proven correct. This allows us to simplify some aspects of the FARM design. In particular, FARM has a more complex way of determining the state of the new leader upon a reconfiguration, which merges the states from all surviving replicas of the previous configuration. In contrast, our protocols take the state of any single initialized replica. Our reconfiguration protocols also provide better fault-tolerance guarantees on a par with those of existing ones [3, 26]. This is because, like Vertical Paxos I [19], our protocols look through a sequence of configurations to find the new leader, whereas FARM only considers the previous configuration. Hence, FARM reconfiguration can get stuck even when there exists a non-faulty replica with the necessary data. Finally, by presenting two related protocols using message passing and RDMA, we are able to identify the price of exploiting RDMA—having to reconfigure the whole system instead of a single shard.

There have been a number of protocols for solving the atomic commit problem, which requires reaching a decision on a single transaction [10, 12, 13, 25]. In contrast to these works, our protocol solves the more general problem of implementing a Transaction Certification Service, which requires reaching decisions on a stream of transactions. This problem more faithfully reflects the requirements of modern transaction processing systems [6].

Our protocol weaves together two-phase commit (2PC) [12] and Vertical Paxos [19], instead of using Paxos replication as a black box. This is similar to several existing sharded systems for transaction processing, which integrate protocols for distribution and replication [6, 16, 22, 28]. However, these systems considered a static set of $2f + 1$ processes per shard, whereas we assume $f + 1$ processes and allow the system to be reconfigured. Achieving this correctly is nontrivial and requires a subtle interplay between the reconfigurable replication mechanism and cross-shard coordination. For example, as we explained in §3, on failures our protocol may lose information about transactions that influenced votes on other transactions, but this does not violate correctness. As is well-known [20], using $f + 1$ instead of $2f + 1$ replicas results in somewhat weaker availability guarantees: upon a single failure, our protocols have to stop processing transactions while the system is reconfigured.

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