Recovering Shared Objects Without Stable Storage

[Extended Version]*

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Abstract

This paper considers the problem of building fault-tolerant shared objects when processes can crash and recover but lose their persistent state on recovery. This Diskless Crash-Recovery (DCR) model matches the way many long-lived systems are built. We show that it presents new challenges, as operations that are recorded at a quorum may not persist after some of the processes in that quorum crash and then recover.

To address this problem, we introduce the notion of *crash-consistent quorums*, where no recoveries happen during the quorum responses. We show that relying on crash-consistent quorums enables a recovery procedure that can recover all operations that successfully finished. Crash-consistent quorums can be easily identified using a mechanism we term the *crash vector*, which tracks the causal relationship between crashes, recoveries, and other operations.

We apply crash-consistent quorums and crash vectors to build two storage primitives. We give a new algorithm for multi-writer, multi-reader atomic registers in the DCR model that guarantees safety under all conditions and termination under a natural condition. It improves on the best prior protocol for this problem by requiring fewer rounds, fewer nodes to participate in the quorum, and a less restrictive liveness condition. We also present a more efficient single-writer, single-reader atomic set—a *virtual stable storage* abstraction. It can be used to lift any existing algorithm from the traditional Crash-Recovery model to the DCR model. We examine a specific application, state machine replication, and show that existing diskless protocols can violate their correctness guarantees, while ours offers a general and correct solution.

1 Introduction

Today's distributed systems are key pieces of infrastructure that must remain available even though the servers that implement them are constantly failing. These systems are long-lived and must be able to tolerate nodes crashing and rejoining the system. In particular, nodes must be able to rejoin the system even after losing their disk state, a real concern for large-scale data centers where hard drive failures are a regular occurrence [7, 36].

*This document is an extended version of the paper by the same title that appeared at DISC 2017 [31]. It contains several appendices which did not appear in the conference proceedings. This document also serves as an updated version of a previous technical report [30].

This paper addresses the problem of how to build recoverable shared objects even when processes lose their entire state. We consider the *Diskless Crash-Recovery* model: each process in the system may go down at any time; upon recovery, it loses all state it had before the crash except for its identity. However, processes can run a *recovery protocol* to reconstruct their state before deeming themselves operational again. This model matches the way that many distributed systems are built in practice.

The Diskless Crash-Recovery model (DCR) is more challenging than the traditional Crash-Stop model (CS) or the Crash-Recovery with Stable Storage model (CRSS). The main challenge is that an invariant that holds at one process may not hold on that process's next incarnation after recovery. This leads to the problem of *unstable quorums*: it is possible for a majority of processes to acknowledge a write operation, and yet processes can still subsequently lose that knowledge after crash and recovery.

We provide a general mechanism for building recoverable shared objects in the DCR model. We show that an operation can be made recoverable once it is stored by a *crash-consistent quorum*, which we informally define as one where no recoveries happen during the quorum responses. Crash-consistent quorums can be efficiently identified using a mechanism called the *crash vector*: a vector, maintained by each process, that tracks the latest known incarnation of each process. By including crash vectors in protocol messages, processes can identify the causal relationship between crash recoveries and other operations. This makes it possible to discard responses that are not part of a crash-consistent quorum. We show that this is sufficient to make storage mechanisms recoverable.

The crash-consistent quorum approach is a general strategy for making storage primitives recoverable. We give two concrete examples in this paper, both of which are always safe and guarantee liveness during periods of stability; other storage primitives are also possible:

- First, we build a *multi-writer, multi-reader atomic register* by extending the well-known ABD protocol [3] with crash vectors. This improves on the best prior protocol by Konwar et al. [20], $RADON_R^{(S)}$, for this problem: it requires fewer rounds (2 rather than 3), requires fewer nodes to participate in the protocol (a simple majority vs ¾), and has a less restrictive liveness condition.

- Second, we construct a *single-writer, single-reader atomic set*, which has weaker semantics yet permits a more efficient implementation, requiring only a single round of communication for writes. We refer to this algorithm as *virtual stable storage*, as it offers consistency semantics similar to a local disk. We show that the virtual stable storage protocol can be used to transform any protocol that operates in the traditional CS or CRSS models to one that operates in DCR.

We discuss the application of this work to state machine replication, a widely used distributed system technique. Recovering from disk failures is an important concern in practice, and recent replication protocols attempt to support recovery after complete loss of state. Surprisingly, we find that each of the three such protocols [7, 19, 25] can lose data. We identify a general problem: while these protocols go to great lengths to ensure that a recovering replica reconstructs the set of operations it previously processed, they fail to recover critical *promises* the replica has previously made, e.g., to elect a new leader. This is due to the fact that these protocols rely on *unstable quorums* to persist these promises. This causes nodes to break important invariants upon recovery, causing the system to violate safety properties. Our approach provides a correct, general, and efficient solution.

To summarize, this paper makes the following contributions:

- It formalizes a *Diskless Crash-Recovery (DCR)* failure model in a way that captures the challenges of long-lived applications (Section 3).

- It introduces the notion of *crash-consistent quorums* and provides two communication primitives for reading from and writing to crash-consistent quorums (Section 4).
- It presents algorithms built on top of our communications primitives for two different shared objects in the DCR model: an atomic multi-writer, multi-reader register and an atomic single-writer, single-reader set. The former is a general purpose register which demonstrates the generality of our approach, while the latter provides a *virtual stable storage* interface that can be used to port any protocol in the CRSS model to one for the DCR model (Section 5).
- Finally, it examines prior protocols for state machine replication in the DCR failure model and demonstrates flaws in these protocols that lead to violations of safety properties. Our two communication primitives can provide correct solutions (Section 6).

2 Background and Related Work

Static Systems. A static system comprises a fixed, finite set of processes. Fault-tolerant protocols for reliable storage for static systems have been studied extensively in the Crash-Stop (CS) failure model, where processes that fail never rejoin the system, and the Crash-Recovery with Stable Storage model (CRSS). In the latter model, processes recover with the same state after a crash. Consensus and related problems, in particular, have been studied extensively in these settings [8, 15, 34]. In CRSS, a crashed and recovered node is no different than one which was temporarily unavailable; asynchronous algorithms that tolerate lossy networks are inherently robust to these types of failures [8].

Our work addresses the challenge of implementing fault-tolerant shared objects in a *Diskless Crash-Recovery (DCR)* model, i.e., without stable storage. In particular, we demonstrate a communication primitive which persists quorum knowledge even in the presence of diskless recoveries, in the spirit of [3, 13].

Prior work on fault-tolerant shared objects and consensus without stable storage generally requires some subset of the processes to never fail [1, 11]. Aguilera et al. [1] showed an impossibility result for a crash-recovery model: even with certain synchrony assumptions, consensus cannot be solved *without at least one process that never crashes*. The main differentiator between that work and this paper is that in their model, the states of processes were binary—either "up" or "down." We overcome this limitation by adding an extra "recovering" state. As long as the number of processes which are "down" or "recovering" at any given time is bounded, certain problems can be solved even *without processes that never fail*. This addition is especially well suited to capturing the needs of long-lived applications in which processes occasionally fail and need to rebuild state.

Recently, Konwar et al. [20] presented a set of algorithms for implementing an atomic multi-writer, multi-reader (MWMR) register in a model similar to ours. We generalize and improve on this work using new primitives for crash-consistent quorums. Our techniques are applicable to other forms of shared objects as well, and our MWMR register is more efficient: it requires one fewer phase and a simple majority quorum (vs ¾). Like that work, we also draw on the ABD protocol [3] for our implementation of an atomic MWMR register. We also draw on the network stability models in [20] to characterize the conditions under which our shared objects guarantee progress.

Recovering without disk state is an important practical concern for real systems. Several recent practical state machine replication systems [7, 19, 25] incorporate ad hoc recovery mechanisms for nodes to recover from total disk loss. The common intuition behind these approaches is that a

synchronous write to disk can be replaced with a write to a quorum of other nodes, recovering after a failure by performing a quorum read. However, we show that these protocols are not correct; they can lose data in certain failure scenarios. A more recent design, Replacement [16], provides a mechanism for replacing failed processes. Like our work and the epoch vectors in JPaxos [19], it draws on concepts like version vectors [35] and vector clocks [9] to determine the causal dependencies between replacements and other operations. We build on these techniques to provide generic communication primitives in DCR.

Dynamic Systems. In a dynamic setting, processes may leave or join the system at will. At any given time, the system membership consists of the processes that have joined and have not yet left the system. Although we consider a static system, DCR may be viewed as a dynamic system with a finite concurrency level [29], i.e, where there is a finite bound on the maximum number of processes that are simultaneously active, over all runs. Here, a recovering process without state is equivalent to a newly joined process, and a process which crashes is equivalent to a process that leaves the system without an explicit announcement .

Many dynamic systems implement *reconfiguration* protocols [2, 10, 24–27, 38]. Reconfiguration allows one to change the set of members allowed to participate in the computation. This process allows both adding new processes and removing processes from the system. Reconfiguration is a more general problem than recovery: it can be used to handle disk failure by introducing a recovering node without state as a new member and removing its previous incarnation. However, general reconfiguration protocols are a blunt instrument, as they must be able to handle completely changing the membership to a disjoint set of processes (potentially even with a different size). As a result, these protocols are costly. Most use consensus to agree on the order of reconfigurations, which delays the processing of concurrent operations [32]. DynaStore [2] is the first proposal which does not require consensus, but reconfigurations can still delay R/W operations [32]. SmartMerge [17] improves on DynaStore by offering a more expressive reconfiguration interface. Recovery is a special case of reconfiguration, where each recovering process replaces, and has the same identity as, a previously crashed process. As a result, it permits more efficient solutions.

Other protocols implement shared registers and other storage primitives in churn-prone systems [4–6, 18]. In these systems, processes are constantly joining and leaving the system, but at a bounded rate. These protocols remain safe only when churn remains within the specified bound, in contrast to our work which is always safe. Most of these protocols also require synchrony assumptions for correctness; our protocols remain correct in a fully asynchronous system. However, under these assumptions they are able to provide liveness guarantees even during constant churn.

3 System Model

We begin by defining our failure model: *Diskless Crash-Recovery* (DCR), a variant of the classic Crash-Recovery model where processes lose their entire state upon crashing.

We consider an asynchronous distributed system which consists of a fixed set of n processes, Π. Each process has a unique name (identifier) of some kind; we assume processes are numbered $1, \ldots, n$ for simplicity. Each process executes a protocol (formally, it is an I/O automaton [28]) while it is up. An execution of a protocol proceeds in discrete time steps, numbered with $\mathbb{N}$, starting at $t = 0$. At each step, at most one process either processes an input action, processes a message, crashes, or

restarts. If it *crashes*, the process stops receiving messages and input actions, loses its state, and is considered DOWN. A process that is DOWN can *restart* and transition back to the UP state. We make the following assumptions about a process that restarts: (1) it knows it is restarting, (2) it knows its unique name and the names of the other processes in the system (i.e., this information survives crashes), and (3) it can obtain an incarnation ID that is distinct from all the ones that it previously obtained. Note that the incarnation ID need only be unique among different incarnations of a specific process, not the entire system. These are reasonable assumptions to make for real-world systems: (1) and (2) are fixed for a given deployment, and (3) can be obtained, for example, from a source of randomness or the local processor clock.

Processes are connected by an asynchronous network. Messages can be duplicated a finite number of times or reordered arbitrarily – but not modified – by the network. We assume that if an incarnation of a process remains UP, sends a message, and an incarnation of the destination process stays UP long enough, that message will eventually be delivered.[1]

The unique incarnation ID makes it possible to distinguish different incarnations of the same process. Without unique incarnation IDs, processes are vulnerable to "replay attacks:"

Theorem 1. *Any state reached by a process that has crashed, restarted, and taken steps without receiving an input action or crashing again will always be reachable by that process.*

Proof. Suppose process p has crashed, restarted, and taken some number of steps without crashing or receiving an input action. That is, suppose that after it restarted, p received some (possibly empty) sequence of messages, $\mathscr{M}$. Because p is an I/O automaton without access to randomness or unique incarnation IDs, anytime p crashes and restarts, it restarts into the exact same state. Furthermore, if p crashes, restarts, and receives the same sequence of messages, $\mathscr{M}$, having been duplicated by the network, p will always end up in the same state. □

A corollary to Theorem 1 is that any protocol in the DCR model without unique incarnation IDs satisfying the safety properties of consensus—or even a simple shared object such as a register—can reach a state from which terminating states are not reachable (i.e., a state of deadlock). If all processes crash and recover as in Theorem 1 before deciding a value or receiving a write, they can always return to this earlier state, so the protocol cannot safely make progress.

For simplicity of exposition, we assume that the incarnation ID increases monotonically. Appendix A explains how to eliminate this requirement.

A restarting process must recover parts of its state. To do so, it runs a distinct *recovery protocol*. This protocol can communicate with other processes to recover previous state. Once the recovery protocol terminates, the process declares recovery complete and resumes execution of its normal protocol. We describe a process that is UP as RECOVERING if it is running its recovery protocol and OPERATIONAL when it is running the initial automaton. A protocol in this model should satisfy *recovery termination*: a recovering process eventually completes recovery and becomes OPERATIONAL, as long as it does not crash again in the meantime. This precludes vacuous solutions where recovering process never again participate in the normal protocol.

Using a separate recovery protocol matches the design of existing protocols like Viewstamped Replication [25]. Importantly, the distinction between RECOVERING and OPERATIONAL makes it

[1] This model is equivalent to one in which the network can drop any message a finite number of times, with the added stipulation that processes resend messages until they are acknowledged.

possible to state failure bounds in terms of the number of OPERATIONAL processes, e.g., that fewer than half of the processes can be either DOWN or RECOVERING at any moment. This circumvents Aguilera et al.'s impossibility result for consensus [1], which does not make such a distinction (i.e., restarting processes are immediately considered OPERATIONAL). We show later how to build a stable storage abstraction that can adapt consensus-based protocols from CRSS to DCR; the resulting protocols are safe, live, and (unlike [1]) do not require certain processes to never crash.

4 Achieving Crash-Consistent Quorums

Making shared objects recoverable in the DCR model requires a new type of quorum to capture the idea of persistent, recoverable knowledge. A simple quorum does not suffice. We demonstrate the problem through a simple straw-man example, and introduce the concepts of *crash-consistent quorums* and *crash vectors* to solve the problem. We use these to build generic quorum communication and recovery primitives.

4.1 Unstable Quorums: Intuition

Consider an intentionally simple example: a fault-tolerant *safe* register that supports a single writer and multiple readers. A safe register [22] is the weakest form of register, as the behavior of READ operations is only defined when there are no concurrent WRITEs. We further constrain the problem by allowing the writer to only ever execute one WRITE operation. That is, the only safety requirement is that once the WRITE completes, all subsequent READs that return must return the value written. READs should always return as long as a majority of processes are OPERATIONAL at any time.

In the Crash-Stop model, a trivial quorum protocol suffices: WRITE(*val*) broadcasts *val* to all processes and waits for acknowledgments from a quorum. Here, we consider majority quorums:

Definition 1. *A quorum Q is a set of processes such that $Q \in \mathcal{Q} = \{Q : Q \in 2^{\Pi} \wedge |Q| > n/2\}$.*

A subsequent READ would then be implemented by reading from a quorum. The quorum intersection property (i.e., $\forall Q_1, Q_2 \in \mathcal{Q} \;\; Q_1 \cap Q_2 \neq \{\}$) guarantees that at least one process will return *val* for a READ that happens after the WRITE. It is easy to extend this protocol to the CRSS model simply by having each process log *val* to disk before replying to a WRITE.

Could we use this same quorum protocol in our DCR model, where processes that crash recover without stable storage, by augmenting it with a recovery protocol that satisfies recovery termination? In fact, for this particular protocol, there is *no* recovery protocol that both guarantees the safety requirement and recovery termination – even if there is a majority of processes which are OPERATIONAL at any instant! In order to tolerate the crashes of a minority of processes and satisfy recovery termination, any recovery protocol must be able to proceed after communicating with only a simple majority of processes. However, if a process crashes in the middle of the WRITE procedure—after acknowledging *val*—it may recover before a majority of processes have received *val*. No recovery procedure that communicates only with this quorum of processes can cause the process to relearn *val*.

We term the resulting situation an *unstable quorum*: the WRITE operation received responses from a quorum, and yet by the time it completes there may no longer exist a majority of processes that know *val*. It is thus possible to form a quorum of processes that either acknowledged *val* but then

lost it during recovery, or never received the write request (delayed by the network). A subsequent READ could fail by reading from such a quorum.

Although this is a simple example, many important systems suffer from precisely this problem of unstable quorums. We show in Section 6 that essentially this scenario can cause three different state machine replication protocols to lose important pieces of state.

4.2 Crash-Consistent Quorums

We can avoid this problem – both for the straw-man problem above and in the general case – by relying not just on simple quorums of responses but *crash-consistent* ones.

Crash Consistency. We informally define a *crash-consistent quorum* to be one where no recoveries of processes in the quorum *happen during* the quorum responses. More precisely:

Definition 2. *Let $\mathscr{E}$ be the set of all events in an execution. A set of events, $E \subseteq \mathscr{E}$, is* crash-consistent *if $\forall e_1, e_2 \in E$ there is no $e_3 \in \mathscr{E}$ that takes place at a later incarnation of the same process as e_1 such that $e_3 \rightarrow e_2$. Here, $\rightarrow$ represents Lamport's happens-before relation [21].*

In Section 4.3, we show how to build recoverable primitives using crash-consistent quorums, in which all quorum replies (i.e. the message send events at a quorum) are crash-consistent.

Crash Vectors. How does a process acquire a crash-consistent quorum of responses? The mechanism that allows us to ensure a crash-consistent quorum is the *crash vector*. This is a vector that contains, for each process, its latest known incarnation ID. Like a version vector, processes attach their crash vector to the relevant protocol messages and use incoming messages to update their crash vector. The crash vector thus tracks the causal relationship between crash recoveries and other operations. When acquiring a quorum on a WRITE operation, we check whether any of the crash vectors are inconsistent with each other, indicating that a recovery may have happened concurrently with one of the responses. We then discard any responses from previous incarnations of the recovering process, ensuring a crash-consistent quorum, and thus avoiding the aforementioned problem.

4.3 Communication Primitives in DCR

We now describe in detail two generic quorum communication primitives, one of which acquires a *crash-consistent quorum*, as well as a generic recovery procedure. These primitives require their users to implement an abstract interface: READ-STATE, which returns a representation of the state of the shared object; UPDATE-STATE, which alters the current state with a specific value; and REBUILD-STATE, which is called during recovery and takes a set of state representations and combines them.

The ACQUIRE-QUORUM primitive writes a value to a crash-consistent quorum (i.e., using the UPDATE-STATE interface) and returns the latest state. The READ-QUORUM primitive returns a *fresh*—but possibly inconsistent—snapshot of the state as maintained at a quorum of processes. If ACQUIRE-QUORUM(val) succeeds, then any subsequent READ-QUORUM will return at least one response from a process that *knows* (i.e., has previously updated its state with) *val*.

[2]This new ACQUIRE message has the same *c* and *val* as before but will have an updated crash vector. Sending it is necessary for liveness.

Algorithm 1 Communications primitives

Permanent Local State:
$n \in \mathbb{N}^+$ ▷ Number of processes
$i \in [1,\ldots,n]$ ▷ Process number

Volatile Local State:
$v \leftarrow [\perp$ **for** $i \in [1,\ldots,n]]$ ▷ Crash vector
$op \leftarrow false$ ▷ Operational flag
$R \leftarrow \{\}$ ▷ Reply set
$c \leftarrow 0$ ▷ Message number

1: **upon** SYSTEM-INITIALIZE
2: $op \leftarrow true$
3: **end upon**

4: **upon** RECOVER
5: $v[i] \leftarrow$ READ-CLOCK
6: $\Sigma \leftarrow$ ACQUIRE-QUORUM$(null)$
7: REBUILD-STATE(Σ)
8: $op \leftarrow true$
9: **end upon**

10: **function** ACQUIRE-QUORUM(val)
11: $R \leftarrow \{\}$
12: $c \leftarrow c+1$
13: $m \leftarrow \langle$ACQUIRE$\rangle$
14: $m.c \leftarrow c$
15: $m.val \leftarrow val$
16: **for all** $j \in [1,\ldots,n]$ **do**
17: SEND-MESSAGE(m, j)
18: **end for**
19: Wait until $|R| > n/2$
20: **return** $\{m.s : m \in R\}$
21: **end function**

22: **function** READ-QUORUM
23: $R \leftarrow \{\}$
24: $c \leftarrow c+1$
25: $m \leftarrow \langle$READ$\rangle$
26: $m.c \leftarrow c$
27: **for all** $j \in [1,\ldots,n]$ **do**
28: SEND-MESSAGE(m, j)
29: **end for**
30: Wait until $|R| > n/2$
31: **return** $\{m.s : m \in R\}$
32: **end function**

33: **function** SEND-MESSAGE(m, j)
34: $m.f \leftarrow i$ ▷ Sender
35: $m.v \leftarrow v$
36: Send m to process j
37: **end function**

38: **upon** receiving $\langle$ACQUIRE$\rangle, m$
39: **guard:** op
40: $v \leftarrow v \sqcup m.v$
41: $m' \leftarrow \langle$ACQUIRE-REP$\rangle$
42: **if** $m.val \neq null$ **then**
43: UPDATE-STATE$(m.val)$
44: **end if**
45: $m'.s \leftarrow$ READ-STATE
46: $m'.c \leftarrow m.c$
47: SEND-MESSAGE$(m', m.f)$
48: **end upon**

49: **upon** receiving $\langle$ACQUIRE-REP$\rangle, m$
50: **guard:** $m.v[i] = v[i] \wedge c = m.c$
51: $v \leftarrow v \sqcup m.v$
52: Add m to R
▷ Discard inconsistent, duplicate replies
53: **while** $\exists m' \in R$ **where**
54: $m'.v[m'.f] < v[m'.f]$ **do**
55: Remove m' from R
56: Resend[2] $\langle$ACQUIRE$\rangle$ message to $m'.f$
57: **end while**
58: **while** $\exists m', m'' \in R$ **where**
59: $m'.f = m''.f \wedge m' \neq m''$ **do**
60: Remove m' from R
61: **end while**
62: **end upon**

63: **upon** receiving $\langle$READ$\rangle, m$
64: **guard:** op
65: $v \leftarrow v \sqcup m.v$
66: $m' \leftarrow \langle$READ-REP$\rangle$
67: $m'.s \leftarrow$ READ-STATE
68: $m'.c \leftarrow m.c$
69: SEND-MESSAGE$(m', m.f)$
70: **end upon**

71: **upon** receiving $\langle$READ-REP$\rangle, m$
72: **guard:** $m.v[i] = v[i] \wedge c = m.c$
73: $v \leftarrow v \sqcup m.v$
74: Add m to R
75: **end upon**

The detailed protocol implementing the two primitives and the recovery procedure is presented as pseudo-code in Algorithm 1. We present the algorithm using a modified I/O automaton notation. In our protocol, **procedures** are input actions that can be invoked at any time (e.g., in a higher

level protocol); **functions** are simple methods; and **upon** clauses specify how processes handle external events (i.e., messages, system initialization, and recovery). We use **guards** to prevent actions from being activated under certain conditions. If the **guard** of a message handler or **procedure** is not satisfied, no action is taken, and the message is not consumed (i.e., it remains in the network undelivered).

Each of the n process in Π maintains a *crash vector*, v, with one entry for each process in the system. Entry i in this vector tracks the latest known incarnation ID of process i. During an incarnation, a process numbers its ACQUIRE and READ messages using the local variable c to match messages with replies. When a process recovers, it gets a new value from its local, monotonic clock and updates its incarnation ID in its own vector. When the recovery procedure ends, the process becomes OPERATIONAL and signals this through the op flag. A process's crash vector is updated whenever a process learns about a newer incarnation of another process. Crash vectors are partially ordered, and a join operation, denoted $\sqcup$, is defined over vectors, where $(v_1 \sqcup v_2)[i] = \max(v_1[i],\ v_2[i])$. Initially, each process's crash vector is $[\bot, \ldots, \bot]$, where $\bot$ is some value smaller than any incarnation ID.

The ACQUIRE-QUORUM function handles both writing values and recovering. ACQUIRE-QUORUM ensures the persistence of both the process's current crash vector—in particular the process's own incarnation ID in the vector—as well as the value to be written, *val*. It provides these guarantees by collecting responses from a quorum of processes and ensuring that those responses are *crash-consistent*. It uses crash vectors to detect when any process that previously replied *could have crashed* and thus could have "forgotten" the written value.

4.4 Correctness

We show that our primitives provide the same safety properties as writing and reading to simple quorums in the Crash-Stop model. First, we formally define quorum knowledge in the DCR context.

Definition 3 (Stable Properties)**.** *A predicate on the history of an incarnation of a process (i.e., the sequence of events it has processed) is a* stable property *if it is monotonic (i.e., X being true of history h implies that X is true of any history with h as a prefix).*

Definition 4. *If stable property X is true of some incarnation of a process, p, we say that incarnation of p* knows *X.*

Definition 5 (Quorum Knowledge)**.** *We say that a quorum Q* knows *stable property X if, for all processes $p \in Q$, one of the following holds: (1) p is* DOWN*, (2) p is* OPERATIONAL *and knows X, or (3) p is* RECOVERING *and either already knows X or will know X if and when it finishes recovery.*

In our analysis of Algorithm 1, we are concerned with knowledge of two types of stable properties: knowledge of values and knowledge of incarnation IDs. An incarnation of a process knows value *val* if it has either executed UPDATE-STATE(val) or executed REBUILD-STATE with an ACQUIRE-REP message in the reply set sent by a process which knew *val*. Knowledge of a process's incarnation ID, i, is the stable property of having an entry in a crash vector for that process greater than or equal to i.

Next, we define crash-consistency on ACQUIRE-REP messages stamped with crash vectors.

Definition 6 (Crash Consistency)**.** *A set of* ACQUIRE-REP *messages R is* crash-consistent *if* $\forall s_1, s_2 \in R\ s_1.v[s_2.f] \leq s_2.v[s_2.f]$.

Note that Definition 6, phrased in terms of crash vectors, is equivalent to the sending events of the ACQUIRE-REP messages being crash-consistent according to Definition 2.

Definition 7 (Quorum Promise)**.** *We say that a crash-consistent set of* ACQUIRE-REP *messages constitutes a* quorum promise *for stable property X if the set of senders of those messages is a quorum, and each sender knew X when it sent the message.*

Definition 8. *If process p sent one of the* ACQUIRE-REP *message belonging to a quorum promise received by some process, we say that p* participated *in that quorum promise.*

The post-condition of the loop on line 53 guarantees the crash-consistency of the reply set by discarding any inconsistent messages; the next loop guarantees that there is at most one message from each process in the reply set. Therefore, the termination of ACQUIRE-QUORUM (line 10) implies that the process has received a quorum promise showing that *val* was written and that every participant had a crash vector greater than or equal to its own vector *when it sent the* ACQUIRE *message*. This implies that whenever a process finishes recovery, it must have received a quorum promise showing that the participants in its recovery had that process's latest incarnation ID in their crash vectors.

Unlike having a stable property, that a process *participated* in a quorum promise holds across failures and recoveries. That is, we say that a process, not a specific incarnation of that process, participated in a quorum promise. Also note that only OPERATIONAL processes ever participate in a quorum promise, guaranteed by the guard on the ACQUIRE message handler.

4.4.1 Safety

Finally, we are ready to state the main safety properties of our generic read/write primitives.

Theorem 2 (Persistence of Quorum Knowledge)**.** *If at time t, some quorum, Q, knows stable property X, then for all times $t' \geq t$, Q knows X.*

Proof. We prove by strong induction on t' that the following invariant, I, holds for all $t' \geq t$: For all p in Q: (1) p is OPERATIONAL and knows X, (2) p is RECOVERING, or (3) p is DOWN. In the base case at time t, Q knows X by assumption, so I holds.

Now, assuming I holds at all times $t' - 1 \geq t$, we show that I holds at time t'. The only step any process $p \in Q$ could take to falsify I is finishing recovery. If recovery began at or before time t, then because Q knew X, p must know X now that it has finished recovering. Otherwise, if it began after time t, then p must have received some set of ACQUIRE-REP messages from a quorum, all of which were sent after time t. By quorum intersection, one of these messages must have come from some process in Q. Call this process q. Since q's ACQUIRE-REP message, m, was sent after time t and before t', by the induction hypothesis, q must have known X when it sent m. Therefore, p must know X upon finishing recovery since it updates its crash vector and rebuilds its state using m.

Since I holds for all times $t' \geq t$, this implies the theorem. □

Theorem 3 (Acquisition of Quorum Knowledge)**.** *If process p receives a quorum promise for stable property X from quorum Q, then Q knows X.*

Proof. We again prove this theorem by (strong) induction, showing that the following invariant, I, holds for all times, t:

1. If a process receives a quorum promise for stable property X from quorum Q, then Q knows X.

2. If process p ever participated in a quorum promise for X at or before time t, and p is OPERATIONAL, then p knows X.

In the base case, I holds vacuously at $t = 0$. We show that if I holds at time $t-1$, it holds at time t:

First, we consider part 1 of I. If p has received a quorum promise, R, from quorum Q for X, then because R is crash-consistent, we know that *at the time they participated in* R no process in Q had participated in the recovery[3] of any later incarnation of any other process in Q than the one that participated in R. If they had, then by the induction hypothesis(which we can apply as their participation happened before time t) , such a process would have known the recovered process's new incarnation ID when it participated in R, and R would not have been crash-consistent.

Given that fact, we will use a secondary induction to show that for all times, t', all of the processes in Q either: (1) haven't yet participated in R, (2) are DOWN, (3) are RECOVERING, or (4) are OPERATIONAL and know X. In the base case, no process in Q has yet participated in R. For the inductive step, note that the only step any process q could take that would falsify our invariant is transitioning from RECOVERING to OPERATIONAL after having participated in R. If q finished recovering, it must have received a quorum promise showing that the senders knew its new incarnation ID. By quorum intersection, at least one of these came from some process $r \in Q$. We already know r couldn't have participated in q's recovery before participating in R. So by the induction hypothesis, r knew X at the time it participated in q's recovery. Because knowledge of values and incarnation IDs is transferred through ACQUIRE-REP messages, q knows X, completing this secondary induction.

Finally, we know that since p has received R at time t, all of the process in Q have already participated in R, so all of the processes in Q are either DOWN, RECOVERING (and will know X upon finishing recovery), or are OPERATIONAL and know X. Therefore, Q knows X, and this completes the proof that part 1 of I holds at time t.

Now, we consider part 2 of I. Suppose, for the sake of contradiction, that p is OPERATIONAL at time t and doesn't know X, but participated in quorum promise R for X at or before time t. Let Q be the set of processes participating in R. Since p does not know X, p must have crashed and recovered since participating in R. Consider p's most recent recovery, and let the quorum promise it received showing that the senders knew p's new incarnation ID (or a greater one) be R'. Let the set of participants in R' be Q'. By quorum intersection, there exists some $r \in Q \cap Q'$.

It must be the case that r participated in R' before R; otherwise by induction, when r participated in R', it would have known X, and then transferred that knowledge to the current incarnation of p (at time t). r couldn't have participated in R before time t, because then by part 2 of I, it would have known p's latest incarnation ID when participating in R, violating the consistency of R. However, r cannot participate in R at or after time t, either. Because p has received a quorum promise for its new incarnation ID at or before time t, by part 1 of I, Q' knows p's new incarnation ID. By Theorem 2, Q' continues to know this at all later times. Because $r \in Q'$, it must know p's incarnation ID, and thus cannot participate in R without violating its crash-consistency. This contradicts the fact that r participates in R and completes the proof that part 2 holds at time t. □

[3]That is, participated in the quorum promise needed by a recovering process, showing that the senders knew the recovering process's new incarnation ID.

Since ACQUIRE-QUORUM(*val*) obtains a quorum promise for *val*, Theorem 3 implies quorum knowledge of *val*, and Theorem 2 shows that that knowledge will persist for all future time, subsequent ACQUIRE-QUORUMs and READ-QUORUMs will get a response from a process which knows *val*.

4.4.2 Liveness

ACQUIRE-QUORUM and READ-QUORUM terminate if there is some quorum of processes that all remain OPERATIONAL for a sufficient period of time.[4] This is easy to see since a writing or recovering process will eventually get an ACQUIRE-REP from each of these OPERATIONAL processes, and those replies must be crash-consistent. Note that the termination of ACQUIRE-QUORUM implies the termination of the recovery procedure, RECOVER. Therefore, the same liveness conditions are required for ACQUIRE-QUORUM and for recovery termination.

We define a sufficient liveness condition, *LC*, below. It is a slightly weaker version of the network stability condition N_2 from [20]: the period in which processes must remain OPERATIONAL is shorter.

Definition 9 (Liveness Condition (*LC*))**.** *Consider a process p executing either the* ACQUIRE-QUORUM *or* READ-QUORUM *function, ϕ, and consider the following statements: (1) There exists a quorum of processes, Q, all of which consume their respective messages sent from ϕ. (2) Every process in Q either (a) remains* OPERATIONAL *during the interval $[T_1, T_2]$, where T_1 is the point in time at which ϕ was invoked and T_2 the earliest point in time at which p completes the consumption of all the responses sent by the processes in Q or (b) becomes* DOWN *and remains* DOWN *during the same interval after p consumed its response. If these two statements are true for* every *invocation of a* ACQUIRE-QUORUM *or* READ-QUORUM *function, then we say that LC is satisfied.*

Our protocol implementing the group communication primitives is live if *LC* is satisfied. For *LC* to be satisfied, it is necessary that at most a minority of processes are DOWN *at any given time*. Otherwise, no process can ever receive replies from a quorum again.[5]

5 Recoverable Shared Objects in DCR

In this section we demonstrate the benefits of our quorum communication primitives for DCR: generality and efficiency. We present protocols for two different shared objects: a multi-writer, multi-reader (MWMR) atomic register (Algorithm 2) and a single-writer, single-reader (SWSR) atomic set(Algorithm 3). In both protocols, READ and WRITE are intended to be invoked serially.

The first protocol implements a shared, fault-tolerant MWMR atomic register in DCR. It is more efficient and has better liveness conditions than prior work. The second protocol implements a weaker abstraction—a shared, fault-tolerant SWSR atomic set. We use this set as a basic storage primitive to provide processes with access to their own *virtual stable storage* (VSS), an emulation of a local disk. This enables easy migration of protocols to DCR.

[4] We assume that the application-provided READ-STATE, UPDATE-STATE, and REBUILD-STATE functions execute entirely locally and do not block.

[5] In fact, this is true if there is ever a majority of processes that are DOWN *or* RECOVERING, unless there is a set of messages currently in the network that will allow some of them to complete recovery.

Algorithm 2 Multi-writer, multi-reader atomic register in Diskless Crash-Recovery

```
     Volatile Local State:
         (t,d) ← (t_0,d_0) ▷ Value of register

 1:  procedure WRITE(d_new)
 2:  guard: op
     ▷ Get latest timestamp
 3:      Σ ← READ-QUORUM
 4:      (t_max,d_max) ← max(Σ)
     ▷ Write value
 5:      t_new ← (t_max.z + 1, i, v[i])
 6:      ACQUIRE-QUORUM((t_new,d_new))
 7:  end procedure

 8:  procedure READ
 9:  guard: op
     ▷ Get latest register value
10:      Σ ← READ-QUORUM
11:      (t_max,d_max) ← max(Σ)
     ▷ Write latest register value
12:      ACQUIRE-QUORUM((t_max,d_max))
13:      return d_max
14:  end procedure

15:  function UPDATE-STATE(val)
16:      if val.t > t then
17:          (t,d) ← val
18:      end if
19:  end function

20:  function READ-STATE
21:      return (t,d)
22:  end function

23:  function REBUILD-STATE(Σ)
24:      (t,d) ← max(Σ)
25:  end function
```

5.1 Multi-writer, Multi-reader Atomic Register

We present a protocol for implementing a fault-tolerant, recoverable multi-writer, multi-reader (MWMR) atomic register in DCR, which guarantees the linearizability of READs and WRITEs [22]. Our protocol is similar to the ABD protocol [3] but augments it with a recovery procedure. Its pseudo-code is presented in Algorithm 2. Timestamps are used for version control, as in the original protocol. A timestamp is defined as a triple $(z,i,v[i])$, where $z \in N$, $i \in [1..n]$ is the ID of the writing process, and $v[i]$ is the incarnation ID of that process. Timestamps are ordered lexicographically. By replacing each quorum write phase in the original protocol with our ACQUIRE-QUORUM function and each quorum read phase with READ-QUORUM, we guarantee that every successful write phase is visible to subsequent read phases, despite concurrent crashes and recoveries, thus preserving safety in DCR. The REBUILD-STATE function reconstructs a value of the register at least as new as the one of the last successful write that finished before the process crashed.

Discussion. The most recent protocol for fault-tolerant, recoverable, MWMR atomic registers is *RADON* [20]. The always-safe version of *RADON*, $RADON_R^{(S)}$, introduces an additional communication phase after each quorum write to check whether any of the processes that acknowledged the write crashed in the meantime. This increases the latency of both the READ and WRITE procedures. Also, our liveness conditions are weaker: our protocol is live if any majority of processes do not crash for a sufficient period of time, while $RADON_R^{(S)}$ requires a supermajority (¾) of processes to not crash.

Algorithm 3 Single writer, single reader atomic set in Diskless Crash-Recovery

Permanent Local State:
owner ▷ Owner of set flag

Volatile Local State:
$S \leftarrow \{\}$ ▷ Local set

1: **procedure** WRITE(s)
2: **guard:** $op \wedge owner$
3: ACQUIRE-QUORUM($\{s\}$)
4: $S \leftarrow S \cup \{s\}$
5: **end procedure**
6: **procedure** READ
7: **guard:** $op \wedge owner$
8: **return** S
9: **end procedure**
10: **function** UPDATE-STATE(val)
11: $S \leftarrow S \cup val$
12: **end function**
13: **function** READ-STATE
14: **return** S
15: **end function**
16: **function** REBUILD-STATE(Σ)
17: ACQUIRE-QUORUM(Σ)
18: $S \leftarrow \bigcup \Sigma$
19: **end function**

5.2 Virtual Stable Storage

Algorithm 3 presents a protocol for a fault-tolerant, recoverable, SWSR set, where the reader is the same as the writer. It guarantees that the values written by completed WRITEs and those returned in READs are returned in subsequent READs. Given the group communication primitives, its implementation is straightforward; the only additional detail is that values read during recovery should be written back to ensure atomicity (line 17).

Discussion. We can use this set to provide a *virtual stable storage* abstraction. It is well known that any correct protocol in CS can be transformed into a correct protocol in the CRSS model by having processes write every message they receive (or the analogous state update) to their local disk before sending a reply. By equipping each process with VSS, any correct protocol in the CRSS model can then be converted into a safe protocol in the DCR model, wherein processes write to crash-consistent quorums instead of stable storage.

This conversion method, while general, might not be the most efficient for a particular problem; more efficient implementations could utilize the communications primitives in Algorithm 1 directly.

6 Recoverable Replicated State Machines in DCR

We further extend our study of DCR to another specific problem: state machine replication (SMR). SMR is a classic approach for building fault-tolerant services [21, 37] that calls for the service to be modeled as a deterministic state machine, replicated over a group of replicas. System correctness requires each replica to execute the same set of operations in the same order, even as replicas and network links fail. This is typically achieved using a consensus-based replication protocol such as Multi-Paxos [23] or Viewstamped Replication [25, 33] to establish a global order of client requests. Once consensus has been achieved, the replicas execute the request and respond to the client.

We examined three diskless recovery protocols for SMR: Viewstamped Replication [25], Paxos Made Live [7], and JPaxos [19]. We found that each of these protocols suffers from the problem illustrated in the example at the beginning of Section 4: they use regular quorums of responses

(instead of crash-consistent ones) when persisting critical data, which could violate their invariants. This can lead to operations being lost, or different operations being executed at different replicas, both serious correctness violations. We provide here brief explanations of the problems in each of these protocols. For more details on how the protocols work and complete traces, see Appendix B.

Viewstamped Replication [33] is the first consensus-based SMR protocol. The original version of the protocol requires a single write to disk, during a view change. A recent VR variant [25] replaces the write to disk with a write to a quorum of replicas, in an attempt to eliminate the necessity for disks. However, it uses simple quorum responses, allowing the recovering replica to violate an important invariant: once a replica committed to take part in a new view, it will never operate in a lower view. As a result, an operation can complete successfully and then be lost after a view change.

Paxos Made Live [7] is Google's Multi-Paxos implementation. To handle corrupted disks, it lets a replica rejoin the system without its previous state and runs an (unspecified) recovery protocol to restore the application state. The replica must then wait to observe a full instance of successful consensus before participating. This successfully prevents the replica from accepting multiple values for the same instance (e.g., one before and one after the crash). However, it does *not* prevent the replica from sending different promises (i.e., leader change commitments) to potential new leaders, which can lead to a new leader deciding a new value for a prior successful instance of consensus.

JPaxos [19], a hybrid of Multi-Paxos and VR, provides a variety of deployment options, including a diskless one. Nodes in JPaxos maintain an *epoch vector* that tracks which nodes have crashed and recovered to discard lost promises made by prior incarnations of recovered nodes. However, like VR and PML, this approach encounters the same problem at a different level: certain failures during node recovery can cause the system to lose state and violate safety properties.

All of these protocols can be correctly migrated to DCR, with little effort, using VSS write operations, as explained in Section 5.2. This approach is straightforward, efficient, and requires no invasive protocol modifications. As an example, we implemented the Viewstamped Replication using VSS. We evaluate this protocol experimentally in Appendix C, showing that it efficiently handles recovery.

7 Conclusion

This paper examined the Diskless Crash-Recovery model, where process can crash and recover but lose their state. We show how to provide persistence guarantees in this model using new quorum primitives that write to and read from *crash-consistent quorums*. These general primitives allow us to construct shared objects in the DCR model. In particular, we show a MWMR atomic register protocol requiring fewer communication rounds and weaker liveness assumptions than the best prior work. We also build a SWSR atomic set that can be used to provide each process with *virtual stable storage*, which can be used to easily migrate any protocol from traditional Crash-Recovery models to DCR.

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A Using Unique IDs Instead of a Monotonic Clock

The communications primitives in Algorithm 1 assume that READ-CLOCK (line 5) returns an integer which is greater than all integers previously returned by READ-CLOCK for that process, across all incarnations. This assumption of monotonicity can be discarded, however, with a slightly more complicated recovery/crash vector mechanism (as long as processes still have the ability to generate *unique* incarnation identifiers, as required by Theorem 1).

There are two approaches: (1) using a crash vector composed of sets of IDs rather than integers, and (2) emulating a monotonic clock (using a nonce for freshness). The former requires that processes keep a crash vector which grows linearly with the number of crashes that have happened in the system, while the latter requires an extra round of communication on recovery.

A.1 Crash Vectors of Sets of IDs

One way to eliminate the monotonicity requirement is to make each process's entry in the crash vector a *set* of unique IDs. The READ-CLOCK call would then be replaced with a call to a function which returns a singleton containing a unique value, i.e., a value never before used by that process. Then, the join operation on these crash vectors would be the element-wise union of the sets. Finally, the guards to the ACQUIRE-REP and READ-REP handlers (lines 50 and 72) would be rewritten as $m.v[i] \geq v[i] \wedge c = m.c$, where $v_1 \leq v_2 \iff \forall i \; v_1[i] \subseteq v_2[i]$.

This modified algorithm works by ensuring that processes write a new, unique value into the crash vectors of a crash-consistent quorum during recovery, so that messages from previous incarnations (i.e., those that don't have the latest unique value) can be ignored. Unfortunately, this means that the size of the crash vectors grows linearly with the total number of recoveries which have happened, instead of with the logarithm of that number.

A.2 Emulating a Monotonic Clock

The other way to forgo monotonicity is to emulate a monotonic clock. The READ-CLOCK call would be replaced with a call to READ-QUORUM to get the latest successfully written crash vector entry for the recovering process (using a nonce on those initial READ messages to guarantee freshness). The recovering process would then increment its entry in the crash vector and continue recovery as normal.

This modification is simpler and more space-efficient than the previous, but it comes with the disadvantage of an extra round of communication during recovery.

B Safety Violations Traces

We show details and full traces of the safety violation appearing in existing SMR proposals for the Diskless Crash-Recovery model.

B.1 Viewstamped Replication

Originally introduced by Oki and Liskov in 1988 [33], Viewstamped Replication was one of the first consensus-based SMR protocols. VR provides linearizability [14] and guarantees liveness as long as a majority of replicas stay up for long enough during the periods of synchrony.

VR is a leader-based algorithm: the system moves through a series of numbered views, in which one node is designated as the leader. VR uses 2 protocols. During normal case execution, the leader assigns sequence numbers to incoming client requests, sends PREPARE messages to replicas, and executes the operation once it has received replies from a majority. When the leader is suspected to have failed, a *view change* protocol replaces the leader with a new one. Replicas increment their view numbers, stop processing requests in the old view, then send the new leader a VIEW-CHANGE message with their log of operations. The new leader begins processing only when it receives VIEW-CHANGE messages from a majority of replicas, ensuring that it knows about all operations successfully completed in prior views. These two protocols are equivalent to the two phases of Paxos.

Correctness for the VR protocol hinges on the following promise: once a node has sent a VIEW-CHANGE message, it processes no further requests from the previous view's leader. The protocol ensures this invariant by having processes increment their view number and only process PREPARE messages with matching view number.

As a result, ensuring that nodes in VR can recover from failures requires that the recovery procedure continue to maintain this promise. That is, *each replica must recover in a view number at least as high as the view number in any* VIEW-CHANGE *message it has ever sent.* The original version of VR [33] achieved this invariant by writing view numbers to stable storage during view changes. A later version, "VR Revisited" [25], claimed to provide a diskless mode of operation. It used a recovery protocol and an extension to the view change protocol to, in essence, replace each write to disk with communication with a quorum of nodes. We show below that this protocol is insufficient to ensure continued correctness of the system.

Recovery Protocol VR Revisited's recovery protocol is straightforward: the recovering replica sends a RECOVERY message to all other replicas.[6] If not recovering or in the middle of a view change, every other node replies with a RECOVERY-RESPONSE containing its view number; the leader also includes its log of operations. Once the recovering replica has received a quorum of responses with the same view number, including one from that view's leader, it updates its state with the information in the log.

VR Revisited adds another phase to the view change protocol. When nodes determine a view change is necessary, they increment their view number, stop processing requests in the old view, and send a START-VIEW-CHANGE message to all other replicas. Only when replicas receive START-VIEW-CHANGE messages from a quorum of replicas do they send their VIEW-CHANGE mes-

[6] This message contains a unique nonce to distinguish responses from different recoveries if a node recovers more than once.

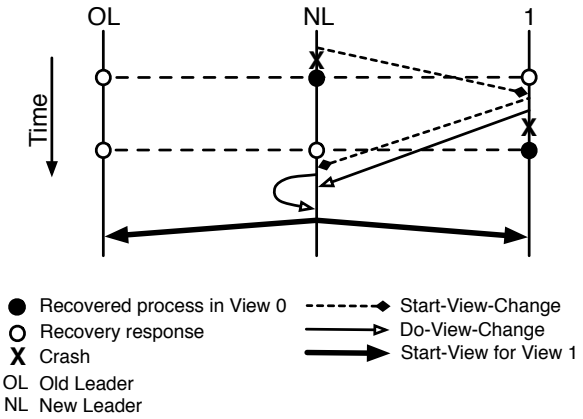


Figure 1: Trace showing safety violation in Viewstamped Replication Revisited [25]

sage to the new leader and proceed as in the original protocol. The additional phase is intended to serve as the equivalent of a disk write, ensuring that replicas do not commit to a new view until a majority of them become aware of the intended view change. Together with the recovery protocol, the added phase aims to prevent violation of the view change invariant by ensuring that a crashed replica recovers in a view at least as high as any VIEW-CHANGE message it has sent.

By itself, however, this approach is not sufficient. The problem of persistence has only shifted a layer: rather than a node "forgetting" that it sent a VIEW-CHANGE message on a crash, it can forget that it sent a START-VIEW-CHANGE.

Failure Trace In Figure 1, we show a trace with 3 replicas, in which a new leader (NL) mistakenly overwrites the decision of a previous leader (OL).

- ⇒ Initially, NL suspects OL of failing and sends a START-VIEW-CHANGE message to node 1 to switch to view 1.
- ⇒ NL crashes immediately after sending this message – before node 1 receives it – then immediately initiates recovery. It sends RECOVERY messages and receives RECOVERY-RESPONSE messages from OL and node 1, both of which are in view 0 – so NL recovers in view 0.
- ⇒ Node 1 receives the START-VIEW-CHANGE message sent by the previous incarnation of NL
- ⇒ Node 1 sends a START-VIEW-CHANGE message to NL for view 1. Because node 1 has a quorum of START-VIEW-CHANGE messages for view 1 (its own and the one from NL), it also sends a DO-VIEW-CHANGE message to NL. Both messages are delayed by the network.
- ⇒ Node 1 crashes and immediately recovers, sending RECOVERY messages to and receiving responses from OL and NL – both of which are in view 0.
- ⇒ NL receives START-VIEW-CHANGE and DO-VIEW-CHANGE messages from node 1. It has a quorum of START-VIEW-CHANGE messages, so it sends a DO-VIEW-CHANGE message for view 1. This is enough for it to complete the view change. It sends a START-VIEW message.
- ⇒ Until the START-VIEW message is received, nodes OL and 1 are still in view 0 and do not believe a view change is in progress. Thus, they can commit new operations, which the new leader NL will not know about, leaving the system in an inconsistent state..

In this trace, messages are reordered inside the network. In particular, messages are reordered

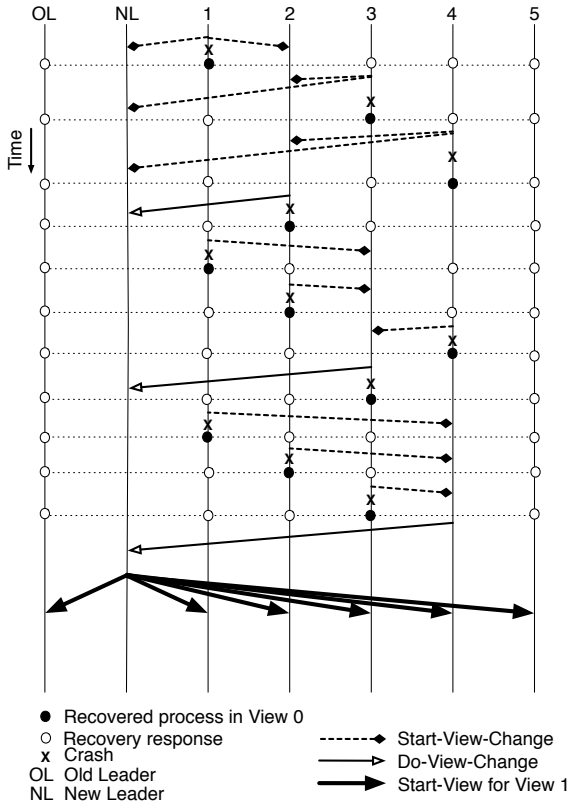


Figure 2: Trace showing safety violation in Viewstamped Replication Revisited [25], even when FIFO channels are assumed

across failures, i.e., messages from a prior incarnation of a node are sometimes delivered *after* messages from a later incarnation. A network with FIFO communication channels may be able to avoid the particular violation described above.

However, message reordering is *not required* in general: there exists a trace with 7 nodes that leads to the same behavior, even with FIFO communication channels, as shown in Figure 2.

B.2 Paxos Made Live

Paxos Made Live is Google's production implementation of a Paxos [7]. It is based on the well-known Multi-Paxos optimization which chains together multiple instances of Paxos [23] and is effectively equivalent to VR. This system primarily uses stable storage to support crash recovery, but because disks can become corrupted or otherwise fail, the authors propose a version that allows recovery without disks.

Recovery Protocol On recovery, a process first uses a *catch-up* mechanism to bring itself up-to-date. The specific mechanism it uses is not described, but presumably it is an application-level state transfer from a quorum of correct processes as in VR. In order to ensure consistency, the recovering process is not allowed to participate in the protocol until it observes a completed instance of successful consensus after its recovery, i.e., until it learns that at least a quorum have agreed on a value for a new

consensus instance. This mechanism suffers from a similar problem to the one in VR. Although it protects against losing ACKNOWLEDGMENTs (i.e., PREPARE-OK messages in VR), it does not protect against losing PROMISEs made to potential new leaders (i.e., VR's DO-VIEW-CHANGE messages).

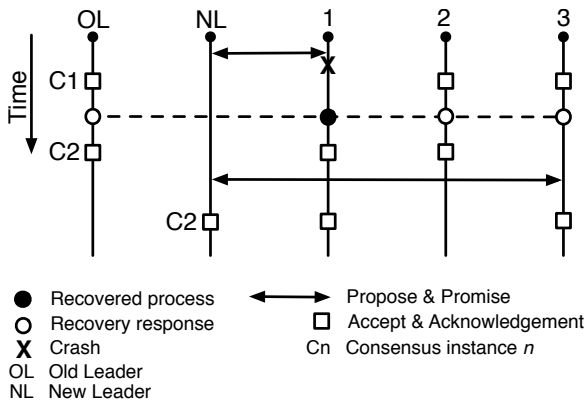


Figure 3: Trace showing safety violation in Paxos Made Live [7]

Failure Trace Figure 3 shows a trace with 5 processes that leads to a new leader mistakenly deciding a new value for a prior successful instance of consensus, overwriting the decision of the previous leader:

⇒ Initially, node OL is the leader.
⇒ NL suspects the leader of having failed and sends a PROPOSE message proposing itself as the next leader.
⇒ Node 1 receives NL's proposal and sends a PROMISE to NL, promising that it will not accept any further operations from OL.
⇒ Node 1 crashes and immediately recovers.
⇒ OL selects a new value for the next instance, C1, and sends ACCEPT and receives ACKNOWLEDGMENT messages from nodes 2 and 3. The operation is committed, and node 1 completes recovery because it has now observed an instance of consensus.
⇒ OL selects a value for instance C2, and sends ACCEPTs and receives ACKNOWLEDGMENTs from nodes 1 and 3.
⇒ Node 3 then receives NL's PROPOSE and sends NL a PROMISE.
⇒ NL has now received a quorum of PROMISE messages: from itself, node 3, and the previous incarnation of node 1. None of these observed consensus instance C2, so NL can now start instance 2 of consensus and overwrite the previous value with its own ACCEPT messages.

B.3 JPaxos

JPaxos [19] is a state machine replication protocol based on Paxos. It is a hybrid between Multi-Paxos and VR, replacing *promises* with *views*. The protocol is presented in several deployment options, including one that does not use stable storage, i.e., supports the Diskless Crash-Recovery

model.[7]

Recovery Protocol JPaxos provides a protocol for the DCR model called *epoch-based recovery*. This protocol is similar in spirit to our use of crash vectors: it attempts to maintain a vector clock of how many times each node has crashed and recovered. Each node updates this vector when it sends a recovery response to another node, and includes a copy of its vector in the PROPOSE-OK messages that it uses to commit to a new view (i.e., the equivalent of the DO-VIEW-CHANGE message). It then uses these vectors to discard all PROPOSE-OK messages that do not come from a crash-consistent quorum.

JPaxos's epoch-based recovery protocol has one critical difference from ours. It treats recovery as complete once it has received RECOVERY-ANSWER messages from a quorum of replicas. In contrast, our protocol requires recovery responses from a *crash-consistent* quorum of replicas. That is, it discards recovery responses from the quorum if it learned that the replica that sent the response itself crashed and recovered.

This subtle change has important implications for correctness. Without this, JPaxos does not correctly handle recovery of nodes when *the nodes they are recovering from* can themselves crash and recover. When this happens, JPaxos can allow a node to recover in a way that simultaneously (1) causes the node to lose the promise that it made in a previous PREPARE-OK, yet (2) does not allow other nodes to detect that the promise might have been lost. We give an example below:

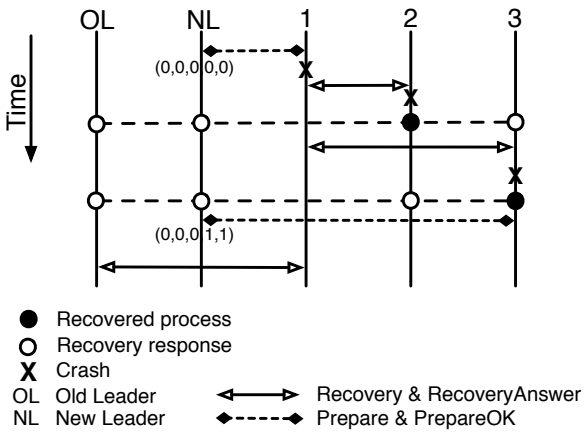


Figure 4: Trace showing safety violation in JPaxos [19]

Failure Trace Figure 4 shows a trace in which a process (node 1 in the figure) commits to a new view via a PREPARE-OK message, then crashes and recovers. When it recovers, it recovers in the previous view without having learned of its promise; however, the new leader (NL) does not learn about node 1's crash and recovery, and therefore does not discard its PREPARE-OK message.

$\Rightarrow$ The system starts in a view where OL is the leader, and all nodes have epoch vector $(0,0,0,0,0)$.

[7]The protocol is actually presented as using stable storage on process initialization, but only to store a monotonic value for each recovery. The monotonic counter available in our DCR model fills this need.

⇒ NL suspects OL of having failed and sends out a PREPARE message proposing itself as the leader of the next view.
⇒ Node 1 receives NL's PREPARE message and sends a PREPARE-OK.
⇒ Node 1 crashes and immediately recovers. It sends a RECOVERY message to node 2 and receives a RECOVERY-ANSWER. Node 2's epoch vector is now $(0,0,1,0,0)$
⇒ Node 2 crashes and immediately recovers. It sends a RECOVERY message and receives RECOVERY-ANSWERs from OL, NL, and node 3. All of these have epoch vector $(0,0,0,1,0)$, so node 2 now has this vector as well – in other words, it has lost its knowledge that node 1 crashed.
⇒ Node 1 sends a RECOVERY message to node 3 and receives a reply. Node 3's epoch vector is now $(0,0,1,1,0)$.
⇒ Node 3 crashes and immediately recovers, communicating with OL, NL, and node 2 during recovery. After recovery, its epoch vector is $(0,0,0,1,1)$.
⇒ NL sends a PREPARE message to node 3 and receives a PREPARE-OK responses. It now has a quorum of PREPARE-OK responses from itself, node 1, and node 3, so it can start a new view.
⇒ Node 1 sends a RECOVERY message to OL, and receives a response. It is now fully recovered in the original view.
⇒ OL can now commit operations (via the quorum of itself, node 1, and node 2) which will not appear in NL's new view.

Our protocol avoids this problem by checking for a crash-consistent quorum on recovery. When node 1 receives a recovery response from OL, that response will have crash vector $(0,0,0,1,1)$ – and so node 1 will discard the earlier recovery responses it received from nodes 2 and 3. It does so because it has learned that those nodes have crashed and recovered, and therefore their updates to the crash vector may not be stable.

C Evaluation of Recovery and Reconfiguration

Virtual stable storage (Section 5.2) is a practical solution for diskless recovery. To demonstrate this, we added an implementation of our recovery protocol to the reference Viewstamped Replication [25] codebase, which provides a general state machine interface. This VR implementation consists of approximately 3500 lines of Java code; less than 100 had to be modified to implement diskless recovery. All clients and replicas ran on servers with 2.5 GHz Intel Xeon E5-2680 processors and 64 GB of RAM. All experiments used three replicas (thereby tolerating one replica failure).

We conducted several experiments to show the effects of our recovery protocol on the throughput of the system. We compared it with the existing recovery protocol from VR, which we showed in Appendix B does not guarantee correctness. Not surprisingly, as our recovery protocol only adds a small amount of metadata to view changes and recovery, the performance is indistinguishable; we do not show these results here.

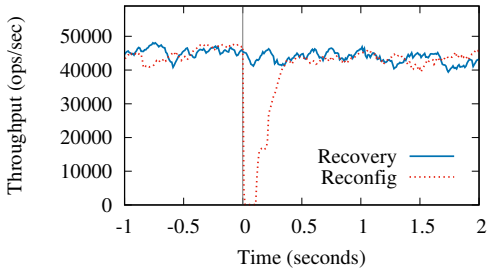
We also compared against the reconfiguration protocol already provided in the VR codebase. The reconfiguration protocol implements the R_1 method described in [24], in which reconfiguration is triggered with a special operation, $rcfg(C)$, and executed using the state machine itself. Method R_1 is more appropriate for recovery purposes than the delayed R_α [24], because processes should be allowed to recover as fast as possible. Our recovery protocol requires just one round trip delay to both persist the new crash vector and recover the lost state.

We ran each experiment multiple times and show the average throughput across those individual experiments to reduce noise from extraneous factors (e.g. Java garbage collection) and better show the impact of recovery and reconfiguration on system throughput.

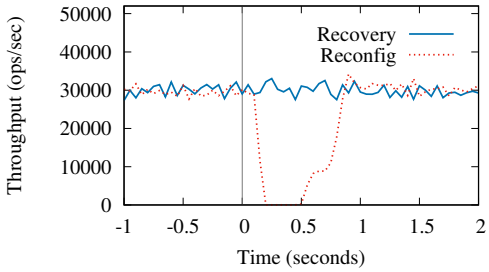
Effect of a single recovery/reconfiguration. We ran our first experiment with cluster-level network latencies ($\approx$ 100 μs) with 24 closed-loop clients, enough to saturate the replicas' load. Figure 5a shows the effect of a single recovery/reconfiguration procedure on the throughput of the system. The blocking nature of the reconfiguration protocol is clearly visible as the throughput drops to 0 ops/sec for at least 100 ms.

Effect of network latency. The second experiment shows the effect of high latency (e.g., in a wide area network) on the throughput during a recovery/reconfiguration. We induced 25 ms additional network latency on our testbed using the Linux traffic control tool, `tc`, emulating a geodistributed deployment. We used 900 closed-loop clients. Our recovery protocol does not affect the overall throughput. In contrast, the impact of reconfiguration is even more visible, as the throughput drops to 0 ops/sec for $\approx$ 300 ms and does not fully recover for nearly a second.

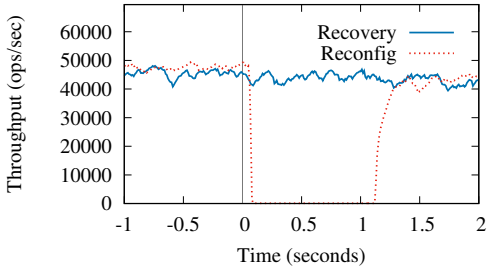
Effect of consecutive recoveries. Another consequence of a blocking reconfiguration protocol is that a single misconfigured node can prevent the system from making progress. We demonstrate this by simulating an unstable process which continuously crashes and restarts, beginning the recovery/reconfiguration procedure anew as soon as the previous recovery/reconfiguration finished. This sort of "flapping" might occur due to a configuration error, or when system load causes slow nodes to be inadvertently suspected of being faulty and removed from the system [12]. As shown in Figure 5c, there is no visible effect on the throughput during consecutive recoveries when no data transfer is conducted (the leader must only process a single message per recovery period). In contrast, multiple consecutive reconfigurations reduced the throughput to 0 ops/sec for the whole period.



(a) Throughput during a single recovery and reconfiguration starting at time 0



(b) Throughput during a single recovery and reconfiguration with wide-area latencies starting at time 0



(c) Throughput during multiple consecutive recoveries and reconfigurations from 0 to 1 second

Figure 5: Experimental results comparing a recovery protocol and a reconfiguration protocol under various scenarios

State transfer. Note that in the preceding experiments, we do not transfer any application data state or the operation log. We do this to simulate the best case scenario for a reconfiguration protocol, which may have to transfer state during the blocking period. Variations of reconfiguration protocols have been proposed that optimize state transfer using various techniques [26]; our experiments assume an ideal, free state transfer. As our recovery protocol is non-blocking, state transfer is done entirely in the background. We have measured system performance for varying amounts of state (transferring a log of up to 500K operations), and state transfer has no impact on system throughput for our recovery protocol.