PAC: Practical Accountability for CCF

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Abstract

Permissioned ledger systems execute transactions on a set of replicas governed by members of a consortium. They use Byzantine fault tolerance protocols to distribute trust among the replicas, and thus can ensure linearizability if fewer than 1/3 of the replicas misbehave. With more misbehaving replicas, current systems provide no guarantees, and all replicas and members share the blame.

We describe PAC, a permissioned ledger system that assigns blame to misbehaving replicas while supporting governance transactions to change the consortium membership and the set of replicas. PAC signs and stores protocol messages in the ledger and provides clients with signed, universally-verifiable receipts as evidence that a transaction executed at a certain ledger position. If clients obtain a sequence of receipts that violate linearizability, anyone can audit the ledger and the sequence of receipts to assign blame to at least 1/3 of the replicas, even if all replicas and members misbehave. Auditing assigns blame by finding contradictory statements signed by the same replica. Since the set of replicas changes, PAC determines the valid signing keys at any point in the ledger using a shorter sub-ledger of governance transactions. PAC provides a strong disincentive to misbehavior at low cost: it can execute more than 48,000 transactions per second, and clients receive receipts in two network round trips.

1 Introduction

Permissioned ledger systems, e.g., Hyperledger Fabric\cite{8}, Quorum\cite{53} and Diem\cite{10}, allow members of a consortium, which are not individually trusted, to deploy a trustworthy service on a set of replicas that they control. They use Byzantine fault tolerant (BFT) state machine replication protocols\cite{18, 24, 27, 32, 43, 61} to distribute trust: clients send requests to execute transactions\cite{58, 59}, the transactions are executed in the same order by the set of replicas, and the results are recorded in a persistent, replicated ledger.

BFT protocols ensure linearizability\cite{36} and liveness but their guarantees only hold if fewer than 1/3 of the $N$ replicas misbehave. If more replicas misbehave, current permissioned ledger systems offer no guarantees and all replicas, and consortium members share the blame equally.

The main approach to avoid this problem has been scaling to more replicas\cite{32, 42, 61}. However, increasing $N$ without adding consortium members does not increase trustworthiness, because replicas do not behave independently if they are controlled by the same member. In real-world deployments, the number of consortium members cannot be increased arbitrarily. For example, the Diem Association\cite{30} has 26 members, and thus cannot offer a service with more than 26 independent replicas—many other consortia are smaller lowering the bar for collusion\cite{6, 7, 15}. Even for large consortia with reputable companies, a patient attacker may compromise $N/3$ replicas over time by exploiting lax security practices, e.g., by bribing employees of consortium members or by exploiting software vulnerabilities. Since there is no accountability, the perceived reputational loss incentivizes members not to disclose these incidents\cite{1, 2, 37}.

Accountability in distributed systems has been explored in the past\cite{8, 33, 34, 62}. Most notably, PeerReview\cite{34} provides accountability in general message passing systems. As we show in \S7, applying this general approach to permissioned ledger systems incurs high overhead: all messages and acknowledgements must be signed, and auditing is expensive because it requires correlating the logs across many replicas. Concurrent work to ours\cite{26, 41} investigates accountability in BFT state machine replication protocols, but it provides no guarantees if 2/3 or more replicas fail. This simplifies the problem because it prevents misbehaving replicas from rewriting the ledger, but leaves these systems vulnerable to a patient adversary who waits until they control at least 2/3 of the replicas before launching an attack. None of these proposals supports changing the replica set or the consortium membership, which is an essential requirement in permissioned ledger systems due to their long-lived nature.

We describe \textit{Practical Accountability for CCF} (PAC), a variant of the CCF\cite{55} permissioned ledger system that can assign blame to misbehaving replicas and the members who control them, \textit{even if} replicas and members change dynamically.
ally and they all misbehave. This strong accountability guarantee improves trustworthiness by creating a strong disincentive for misbehavior.

PAC uses *Ledger PBFT* (L-PBFT), a new BFT state machine replication protocol for partial-synchrony [31] that orders transactions, and stores them in the ledger together with protocol messages from replicas that justify the chosen order. L-PBFT maintains a Merkle tree [47] over the ledger, and includes the root of the tree in protocol messages. Since protocol messages are signed by the replicas, this signature commits them to the entire contents of the ledger.

PAC then issues *receipts* to clients that provide succinct, universally-verifiable evidence that a transaction executed at a given position in the ledger. Receipts include signed protocol messages from multiple replicas that executed the transaction, thus binding them to the prefix of the ledger.

If clients obtain a sequence of receipts that violate linearizability, anyone can audit the ledger and receipts to assign blame to at least $N/3$ replicas. Auditing produces an irrefutable *universal proof-of-misbehavior* (uPoM) in the form of contradictory statements signed by the same replica. The uPoM can then be used by an *enforcer*, e.g., a court, to punish the members responsible for the misbehaving replicas. Since PAC provides accountability even if all replicas and members misbehave, the enforcer may have to compel members to produce a ledger for auditing and to impose sanctions if they fail to do so within a reasonable timeout. This introduces a weak synchrony assumption, but the enforcer chooses a conservative timeout to make blaming correct members unlikely.

As an example of auditing, a client Alice may have a receipt for a transaction that executed at index $i$ in the ledger and that deposited a million dollars into client Bob’s account. If Bob obtains the receipt from Alice and another receipt for a balance query transaction executed at index $j$, $j > i$ that does not show the balance, he may conduct an audit: he engages an enforcer to obtain the relevant ledger fragment, and replays the transactions between $i$ and $j$ to check for consistency with the receipts. If Bob is right, auditing produces a uPoM for at least $N/3$ replicas, which Bob sends to the enforcer to punish consortium members.

PAC supports *governance transactions* to change the set of replicas and consortium members. This complicates receipt verification and auditing because it changes the signing keys that must be considered. PAC therefore records governance transactions in the ledger, which allows clients, replicas and auditors to determine the set of valid signing keys at any point. Clients do not need to keep the full ledger, but only receipts for governance transactions. Since governance transactions are rare, this *governance sub-ledger* is significantly smaller than the full ledger.

Our prototype implementation of PAC provides strong accountability guarantees while achieving high throughput and low latency: it leverages Merkle trees and a commitment scheme to enable a single signature per replica per transac-

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**Fig. 1: Permissioned PAC ledger system**

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2 Overview of PAC

Fig. 1 shows the PAC design and its main components. A deployment of PAC provides a service to a set of *clients*, which are identified by their signing keys. Clients send requests to execute multi-key transactions by calling stored procedures that define the application logic of the service. Transactions are executed by *replicas* against a strictly-serializable key-value store that supports roll-backs at transaction granularity. Each transaction request $t$ reads and/or writes multiple key-value pairs and produces a transaction result $o$.

*Consortium members* are the owners of the service. They are also identified by their signing keys, and may be added or removed over the lifetime of the service. Members can issue *governance transactions* to change the set of consortium members, add or remove replicas, or update stored procedures. The first governance transaction, referred to as the *genesis transaction* ($gt$), establishes the initial set of members and replicas. Its hash is the identity of the service.

**Ledger PBFT** (L-PBFT) is a BFT state machine replication protocol used by replicas to order transaction requests and results. Since L-PBFT is based on PBFT [24], it provides linearizability and liveness if at most $f = [N/3] - 1$ of $N$ replicas fail in a partially-synchronous environment [31].
**Ledger.** L-PBFT maintains an append-only ledger that stores each transaction request \( t \) and the corresponding result \( o \) at an index \( i \). The ledger also includes periodic checkpoints \( cp \) of the key-value store. The key-value store state can thus be reconstructed by replaying the ledger from a checkpoint. Since the consortium membership and thus the replica set is dynamic in PAC, governance transactions are recorded in the ledger. They form a *governance sub-ledger*, which can be used to learn the set of active replicas and members (and their public signing keys) at any ledger index \( i \).

To assign blame, the ledger includes evidence that a batch of transactions was committed by a quorum of replicas. This evidence consists of at least \( N-f \) signed L-PBFT protocol messages for a batch. It allows reasoning about the protocol execution to produce proofs-of-misbehavior for at least \( f+1 \) misbehaving replicas.

All entries in the ledger are leaves of a Merkle tree \( M \) that binds them together. Protocol messages for a transaction batch contain the root of the Merkle tree \( M_i \), after the batch executes. This commits replicas to the whole ledger while allowing succinct proofs that entries are present.

**Receipts** are created by replicas and returned to clients. A receipt \( ⟨ i, i, o ⟩ \) for a transaction request \( t \) states that \( t \) was executed at ledger index \( i \) and produced result \( o \). A receipt consists of \( N-f \) protocol messages for \( t \)'s batch signed by different replicas and a path from the Merkle tree root \( M \) to the leaf that contains an entry for request \( ⟨ t \rangle \) and \( ⟨ i, o ⟩ \). A receipt thus binds the execution of a request to consortium members via the replicas’ signatures over a root of a Merkle tree that contains the executed request. If two or more receipts are inconsistent with any linearizable execution, then at least \( f+1 \) replicas signed contradictory statements and can be consequentially assigned blame.

A novel commitment scheme enables replicas to sign each batch only once by including \( N-f \) commitment nonces in a receipt: a replica generates a nonce for each batch, signs its hash, and sends the unhashed nonce in place of a second signature to confirm execution. The receipt is validated by examining the signature and the hashed/unhashed nonces.

To establish the validity of a receipt, a client must check its signatures based on the set of correct signing keys from the governance sub-ledger. A receipt therefore includes the ledger index of the last governance transaction, and a client must obtain the receipt of this governance transaction and all those preceding it. Clients cache receipts of governance transactions and can fetch missing ones from the replicas.

**Auditing** produces a *universal proof-of-misbehavior* (uPoM) if clients obtain receipts inconsistent with a linearizable execution. PAC’s ledger is universally-verifiable. Any entity, e.g., a client or an external organisation, can act as an auditor: replaying the ledger, checking consistency with receipts, and potentially generating a uPoM. Since all consortium members and replicas may misbehave, an enforcer, e.g., a court, compels replicas to produce a copy of their ledger for auditing and sanctions non-compliance. The enforcer also punishes consortium members based on uPoMs. Clients and replicas do not to trust each other but they trust the enforcer to apply punishment accurately.

After a client passes a sequence of receipts and the corresponding governance sub-ledger to the auditor, the auditor confirms the receipts’ validity by calculating the Merkle tree roots and verifying the replica signatures and nonces. Then they request the enforcer to fetch the ledger fragment corresponding to the receipts from the replicas. The auditor then checks the validity of the checkpoint \( cp \) referenced by the oldest receipt. It replays the ledger from \( cp \) by re-executing transaction requests while also checking for consistency with receipts (including governance transaction receipts). If an inconsistency is found at index \( i \), the auditor creates a uPoM \( ⟨ i, f, cp, R ⟩ \) with a ledger fragment \( f \), the checkpoint \( cp \), and the inconsistent receipt \( R \). The size of \( f \) is bounded by the maximum number of transactions between consecutive checkpoints. The uPoM is then forwarded to the enforcer, which may impose a penalty on the responsible consortium members.

### 3 Ledger replication and receipts

First, we describe how L-PBFT maintains a ledger with transactions and evidence to justify the execution order (§3.1), and how this evidence is used to create transaction receipts for assigning blame (§3.2). We then describe performance optimizations (§3.3). Note that, in this section, we assume a fixed replica set, and we revisit this assumption in §5.

#### 3.1 Ledger PBFT

**Consensus.** Replicas in L-PBFT, as in PBFT, go through a succession of views. A view is identified by a view number \( v \) that determines the primary replica \( p (p = v \mod N) \). As shown in Fig. 2, the primary orders batches of transaction requests received from clients to amortize protocol overhead across all transactions in the batch. It sends a pre-prepare message to the backup replicas, proposing that the batch be ordered at sequence number \( s \) in view \( v \). If the backups accept this proposal, they send prepare messages to all replicas.

A batch is said to have prepared at view \( v \) with sequence number \( s \) if a replica receives a pre-prepare and \( N-f-1 \) matching prepare messages for \( s \) and \( v \), and all batches with lower sequence numbers have also prepared. Unlike PBFT, L-PBFT prepares batches in order and executes requests before

![Fig. 2: L-PBFT protocol with early execution and receipts](image-url)
 Alg. 1: Ledger Practical Byzantine Fault Tolerance (L-PBFT)

| Line | Description |
|------|-------------|
| 1 | on receiveTransactionRequest (t = (request, a, c, H(g(t)), σ_a)) |
| 2 | | Pre: verify(t) |
| 3 | | T ← T ∪ [t] |
| 4 | on sendPrePrepare() |
| 5 | | Pre: isPrimary() ∧ ready ∧ |T| > 0 ∧ hasEvidence ((M, v, x, P)) |
| 6 | | B ← [] |
| 7 | | foreach i ∈ T do |
| 8 | | | B ← B ∪ H((t, i)); (t, i) ← execute(kv, t); L ← L ∥ (t, i, o) |
| 9 | | | E_{r, p, x, s, k} ← getEvidence ((M, v, y, P)) |
| 10 | | | L ← L ∥ (P, x, E_{r, p, x}) |
| 11 | | | X_{v, s} ← createNonce () ; M = getMerkleRoot (L) |
| 12 | | pp = (pre-prepare, v, x, M, H(X_{v, s}), E_{r, p}, nid, r) |
| 13 | | L ← L ∥ pp |
| 14 | | M ← M ∪ {pp}; T ← []; x ← x + 1 |
| 15 | | sendToAllReplicas (pp || B) |
| 16 | on receivePrePrepare (pp = (pre-prepare, v, x, M, H(k), E_{r, p}, σ_r, B)) |
| 17 | | Pre: isBackup() ∧ verify(pp) ∧ ready ∧ x' = x + 1 ∧ X_{v, s} = nil ∧ hasRequests (T, B) ∧ hasEvidence ((M, x', P, E_{r, p})) |
| 18 | | x ← x + 1; M ← M || {pp} |
| 19 | | foreach i ∈ B do |
| 20 | | | (t, i) ← execute(kv, t); L ← L ∥ (t, i, o) |
| 21 | | | E_{r, p, x, s, k} ← getEvidence ((M, v, x, P, E_{r, p})) |
| 22 | | | L ← L ∥ (P, x, E_{r, p}) |
| 23 | | | if getMerkleRoot (L) = nil then |
| 24 | | | undo(pp, s, B, T, [T, {T}]); return |
| 25 | | | L ← L ∥ pp |
| 26 | | | X_{v, s} ← createNonce () |
| 27 | | | p = (prepare, rid, H(X_{v, s}), H(pp), σ_r) |
| 28 | | | sendToAllReplicas (p || {M}); M = M || {p} |
| 29 | on receivePrepare (p = (prepare, b, H(k)), H(pp)) σ_b |
| 30 | | Pre: verify(p) |
| 31 | | M ← M ∪ {p} |
| 32 | on receiveCommit (c = (commit, v, x', r, k)) |
| 33 | | Pre: verify(c) |
| 34 | | M ← M ∪ {c} |
| 35 | on batchPrepared (pp = (pre-prepare, v, x', M', H(k'), E_{r, p}, σ_r)) |
| 36 | | Pre: prepared(pp; M) ∨ exists prepare, rid, H(X_{v, x'}), H(pp), σ_r) in M |
| 37 | | c = (commit, v, r, X_{v, x'}) |
| 38 | | sendToAllReplicas (c); M = M ∪ {c} |
| 39 | onReset (t, i, o) includetx (T, v, x') |
| 40 | | sendReplyToClient (t, (reply, v, x', rid, H(pp), σ_r, X_{v, x'})) |
| 41 | | if shouldSendReceipt (rid, t) then |
| 42 | | | S ← getMerklePath (L, M, t) |
| 43 | | | sendReceiptToClient (t, (reply, v, M, H(k), E_{r, p}, H(t, i, o, S)) |
| 44 | on sendViewChange () |
| 45 | | Pre: currentAppearsFaulty (v) |
| 46 | | | PP = getPLastPrepared (msgs(L) ∪ M) |
| 47 | | | v = v + 1; ready ← true; wc = (view-change, v, rid, PP, nid) |
| 48 | | | sendToAllReplicas (wc); M = M ∪ {wc} |
| 49 | on receiveViewChange (wc = (view-change, v', t, PP, σ_r)) |
| 50 | | Pre: |v'| = v ∧ verify(wc) ∧ hasPrepares (msgs(L) ∪ M ∗ getLast (PP)) |
| 51 | | | M = M ∪ {wc} |
| 52 | | if getViewChanges (M, v') > f ∧ v' > v then |
| 53 | | | v = v'; ready ← false |
| 54 | on sendNewView (v) |
| 55 | | Pre: isPrimary () ∧ ready ∧ hasGetViewChanges (M, v) > N - f |
| 56 | | | (M, E_{r, h, h}, PP_m) = processViewChanges (getViewChanges (M, v)) |
| 57 | | | mv = (view-change, v, M, E_{r, h, h}, nid); L = L ∥ mv |
| 58 | | | sendToAllReplicas (mv); resendPreparesInNewView (PP_m); ready ← true |
| 59 | on receiveNewView (mv = (view-change, v, M, E_{r, h, h}, PP_m)) |
| 60 | | Pre: isPrimary () ∧ ready ∧ hasRequests (T, PP_m) ∧ hasEvidence ((M, PP_m)) |
| 61 | | | ∧ r ≠ rid ∧ ready ∧ hasGetViewChanges (M, E_{r, h, h}) > N - f |
| 62 | | | (M', PP_m') = processViewChanges (getViewChanges (M, E_{r, h, h})) |
| 63 | | if M' = M then |
| 64 | | | L ← L ∥ mv |
| 65 | | | if ready ← processPreparesInNewView (PP_m', PP_m) then |
| 66 | | | return |
| 67 | | | undo(mv, r, M, L) |

Fig. 3: Ledger with evidence and Merkle tree

they prepare. Replicas exchange commit messages when a batch prepares. The batch is committed at sequence number s if it has been prepared by N − f replicas in the same view.

**View changes** are used to change the primary, e.g., due to failure. Each replica sends a view-change message with information about prepared requests. The primary for a new view v' sends a new-view message backed by N − f view-change messages for v'. For each sequence number with a prepared batch in the view-change messages, the primary picks the batch that prepared with the largest view and proposes it in v'. Since all committed requests have also prepared, this ensures linearizability with batch execution ordered by the sequence numbers at which batches commit.

**Maintaining the ledger.** The ledger not only stores committed transactions but also includes evidence in the form of pre-prepare, prepare, commit, view-change, and new-view messages that justify the execution order chosen by L-PBFT.

Fig. 3 shows the entries in the ledger and the Merkle tree that binds them together. Each batch of transactions has a sequence number s that defines a total order for batches in the ledger. For each transaction in the batch, L-PBFT adds an entry to the ledger in the order executed. For example, the entry for T_i has the form (t, i, o) where t is a transaction request, i indicates that the transaction was the ith to execute since the start of the ledger, and o is the result of executing the transaction; pp_{s-1} is the pre-prepare for s − 1, and P_{s-1} and X_{s-1} are evidence that the batch at sequence number s − P committed. (The last entry for batch s is the pre-prepare, not added yet in the figure.) The evidence of commitment must lag behind s because it is not available when the primary sends the pre-prepare for s. It lags by P because L-PBFT pipelines the ordering of up to P ≥ 1 concurrent batches to improve performance.

Alg 1 defines the pseudocode for the protocol. The replica state includes: the current view v and batch sequence number s; a set of transaction requests waiting to be ordered T; a message store M; a nonce store X; a boolean ready indicating whether the replica can send/accept pre-prepare messages; a replica identifier rid; the key-value store kv; and the ledger L.
stored procedure and its arguments, c is the client identifier, 
$H(gt)$ is the hash of the genesis transaction, and $\sigma_i$ is the 
client signature. The client signature and $H(gt)$ ensure that 
requests cannot be forged or moved to a different ledger.

In sendPrePrepare, the primary collects a batch of trans-
action requests, executes them, and appends them to the 
ledger (line 8). Then, it gets evidence $E_{s-P}$ and $K_{s-P}$ of 
commitment for the batch at $s-P$ and appends it to the ledger. 
$E_{s-P}$ is a bitmap indicating the replicas that supplied evi-
dence for the commitment. Fig. 3 shows the ledger after this 
step. Next, the primary creates the pre-prepare message with 
the hash of a fresh nonce $K[s,v]$ and the root of the Merkle 
tree $M_r$, and signs it. The complete message has two ad-
ditional fields, explained below: $I_g$, which is the index of the last 
governance transaction, and $dC$, which is a digest of the state 
of the key-value store at the last checkpoint. $I_g$ allows clients 
to verify receipts with a changing set of replicas (§5.2), and 
dC enables incremental auditing from a checkpoint without 
replaying the ledger from the start (§4). Moving execution 
from the commit phase, as in PBFT, allows L-PBFT to include 
the execution result in the batch’s Merkle tree root, which we 
refer to as early execution.

The nonce $K[s,v]$ is revealed to clients and other repli-
cas only when the primary prepares the batch (lines 38, 40). 
Therefore, having a pre-prepare and the corresponding nonce 
can prove to a third party that the primary prepared the batch 
at $v$ and $s$. The backups also commit to nonces in prepare 
messages (line 27) and reveal them when they prepare the 
batch. This simple nonce commitment scheme allows replicas 
to return replies to clients in two message roundtrips without 
signing replies or commit messages (see Fig. 2).

By signing $M_r$, the primary commits to the entire contents 
of the ledger, including the commitment evidence for $s-P$ that 
it collected and added to the ledger. $E_{s-P}$ contains $N-f-1$ 
prepare messages for sequence number $s-P$ and view $v$ that 
match the pre-prepare at sequence number $s-P$ in the ledger. 
$K_{s-P}$ are the $N-f$ nonces with hashes in the pre-prepare 
and prepare messages in $E_{s-P}$. This evidence is sufficient 
to prove to a third party that the batch at $s-P$ prepared at 
$N-f$ replicas and therefore committed with $s$. It is important 
for the primary to order the evidence to ensure that replicas 
agree on the ledger. If replicas added their own evidence to 
the ledger when they received prepare and commit messages, 
their ledgers could diverge.

The primary sends the pre-prepare to all replicas together 
with a list $B$ of the hashes of the transaction requests in the 
execution order (line 15). The requests are transmitted sepa-
rately by the clients, and the commitment evidence for $s-P$ is 
not included in the message. The pre-prepare messages in L-
PBFT are $O(N)$ in size but the constant is small. We currently 
use 8 bytes in the $E_{s-P}$ bitmap to support up to 64 replicas, 
making pre-prepare messages effectively $O(1)$.

When a backup replica receives the pre-prepare, it rejects 
the message if it already sent a prepare for the same view 
and sequence number ($K[v,s] \neq \text{nil}$). Otherwise, it checks if 
it already has the requests and commitment evidence re-
ferred to by the pre-prepare. Replicas store received transaction 
requests, prepare, and commit messages in non-volatile storage 
(lines 3, 28, 31, 34, and 38) until they receive (or send) a 
pre-prepare that references them. If the backup is missing any 
messages, it requests the primary to retransmit them because 
a correct primary is guaranteed to have them.

After the backup obtains any missing messages, it exec-
utes the requests in the order prescribed by the primary, and 
adds the resulting transaction entries to the ledger (line 20). 
Then, it adds the same $E_{s-P}$ and $K_{s-P}$ as the primary to the 
ledger. At this point, the ledger at the backup should 
be identical to the one at the primary just before the pre-
prepare message is added. The backup checks that the root 
of its Merkle tree matches $M_r$ in the pre-prepare (line 23). 
If not, the message is rejected, the entries for batch $s$ are re-
moved from the ledger, and the transactions are rolled back. 
Otherwise, the backup adds the pre-prepare to the ledger 
and sends a matching prepare messages with the format 
$\langle \text{prepare}, \text{rid}, H(K[v,s]), H(pp)\rangle_{\sigma_{id}}$, where $H(K[v,s])$ com-
mits a fresh nonce, and $H(pp)$ is the pre-prepare’s hash.

L-PBFT ensures deterministic transaction execution by 
agreeing on non-deterministic inputs [25], unless there is a 
bug in the service code. The check in line 23 ensures that such 
bugs can at most affect liveness. If the check fails, the trans-
action requests will be resent by the clients and re-executed 
which can often resolve the non-determinism.

**View changes.** If the primary for view $v$ appears faulty or 
slow, replicas send a view-change message to all other repli-
cas of the form $(\text{view-change}, v+1, \text{rid}, \text{PP})_{\sigma_{id}}$ where $\text{PP}$ is 
a set with the last $P$ pre-prepare messages that prepared locally 
(line 48). Only the last message in $\text{PP}$ is required to 
provide linearizability because it includes the Merkle tree 
root that determines the contents of the ledger up to that point. 
The other messages allow the auditor to verify that replicas in-
clude all prepared but not committed batches in view-change 
messages without their ledger (§4). Their ledgers may no 
longer be available for auditing, e.g., because they have left 
the service.

When replicas receive a view-change message (line 49), 
they fetch missing prepare messages from the sender to prove 
that the last pre-prepare in $\text{PP}$ has prepared. They do not 
accept the view-change message before that. When replicas 
increment $v$, they set $\text{set to false}$ (lines 47, 53) to ensure 
that they do not send or accept pre-prepare messages until 
they completed the new-view successfully (lines 59, 65). 

After accepting $N-f$ view-change messages for the new 
view (line 55), the new primary calls processViewChanges. 
This function picks the view-change message, $vc_{vp}$, with the 
last prepared pre-prepare message, $pp_{vp}$, from those with 
the largest view number. Then it updates its ledger to match 
the Merkle root in $pp_{vp}$ by fetching any missing ledger entries 
from replicas that sent matching prepare messages. Since at
least \(f+1\) of those are correct, this is always possible. The primary checks that all the messages in \(\mathcal{PP}\) of \(v_{\mathcal{PP}}\) appear at the right positions in the ledger; if not, it discards \(v_{\mathcal{PP}}\) and tries again (not shown in Alg. 1).

Next the primary resets the ledger to \(s_{\mathcal{PP}}-P\), because the batches up to this point are guaranteed to have committed, but it saves all the request batches and commitment evidence for sequence numbers between \(s_{\mathcal{PP}}-P\) and \(s_{\mathcal{PP}}\) and returns it in \(\mathcal{PP}_{ov}\). This is needed to resend \(\text{pre-prepare}\) messages for the prepared batches in the new view. The function ends by adding an entry with the \(N-f\) view-change messages that it accepted to the ledger in order of increasing replica identifier; \(h_{vc}\) is the hash of that entry and \(E_{vc}\) is a bitmap with the replicas that sent the messages. It returns the root of the Merkle tree \(M_r\) at this point, \(E_{vc}, h_{vc}\), and \(\mathcal{PP}_{ov}\). The primary appends the new-view to the ledger (line 57), sends it to all replicas, resends the prepared batches in \(\text{pre-prepare}\) messages in the new view, and adds them to the ledger.

When backups receive the new-view, they obtain any missing view-change messages, requests and evidence that it references (line 61), and call processViewChanges. If it returns a Merkle tree root equal to the one in new-view, they accept the message, add it to the ledger and process the \(\text{pre-prepare}\) messages \(\mathcal{PP}_{ov}'\). If these match the batches and evidence in \(\mathcal{PP}_{ov}\) for the same sequence numbers, they are added to the ledger; otherwise, all changes are undone.

### 3.2 Receipts

To enable replicas to produce a single signature per batch, PAC uses the Merkle tree to bind signatures in \(\text{pre-prepare}\) and \(\text{prepare}\) messages to the receipt statements for each transaction in the batch; otherwise, replicas would sign each transaction’s receipt separately introducing a high overhead.

**Creating receipts.** When a batch of transactions described by \(\text{pre-prepare} pp\) prepares at a replica \(r\), view \(v\) and sequence number \(s'\), it sends \(\langle \text{reply}, v, s', \text{rid}, H(pp_i), \sigma_{id}, \mathcal{K}_r[v,s']\rangle\) to every client with a transaction in the batch (Alg. 1, line 40). If the client has several transactions in the batch, only one reply is sent per replica. By revealing the nonces, the replicas provide the client with proof that they claimed to have prepared the batch without requiring a signed reply.

Only a designated replica, which is chosen based on \(t\), sends back the result and the rest of the receipt to the client (line 41). The replica computes a list of sibling hashes \(S\) along the path from the leaf for the transaction to the root of the Merkle tree. For the example of \(T_i\) in Fig. 3, \(S\) would consist of the digest of \(T_{i+1}\), \(M_3\), and \(M_2\), which is sufficient to recompute \(M_r\) given \(T_i\). Then it sends the client \(\langle \text{reply}, v, s', M_r, H(k_p), E_{r-p}, i_p, d_C, H(t), i, o, S\rangle\), where \(i_p\) and \(d_C\) are used for auditing (not shown in Alg. 1), \(i\) is the index where \(t\) executed, and \(o\) is the transaction result.

**Verifying receipts.** The client waits for \(N-f\) replicas to send matching \(\text{reply}\) messages with the same \(v, s\), and \(H(pp_i)\), and for a \(\text{reply}\) message with the same \(v\) and \(s\). Then it recreates the \(\text{pre-prepare}\) and \(\text{prepare}\) messages (with the information in \(\text{reply}\) and the hashes of the nonces) and verifies the signatures and \(H(pp_i)\). (We describe how to determine \(N\) and verify signatures with governance in §5.2.) This step is shared across all transaction requests that the client may have sent in the batch. If this check succeeds, the client verifies \(S\) for each transaction request \(t\). In the example in Fig. 3, the client checks if \(M_r = H(H(M_3)||H(H(t, i, o))||H(T_{i+1}))||M_2)\). If the hashes match, the client has a valid receipt, i.e., a statement signed by \(N-f\) replicas that a transaction request \(t\) executed at index \(i\) and produced a response \(o\); otherwise (or if the client does not receive replies before a timeout), it retransmits the request, selecting a different designated replica to send back reply. (The application is responsible for ensuring exactly-once semantics if needed.)

Clients store the receipt for \(t, i, o\) compactly as \(\langle v, s, H(k_p), E_{r-p}, i_p, d_C, H(pp_i), \sigma_{id}, \mathcal{K}_r[v, s]\rangle\) where \(\mathcal{K}_r[v, s]\) is a list of the signatures in \(\text{pre-prepare}\) messages, \(\mathcal{K}_r\) is a list of nonces, and \(E_{r-p}\) is a bitmap indicating the replicas with entries in \(\mathcal{K}_r\), and \(\mathcal{K}_r\) which are both sorted in increasing order of replica identifier. All receipt components but \(S\) are shared across multiple requests in the same batch.

Clients should store the receipts together with the transaction request and the corresponding response to resolve future disputes. This is not a burden because receipts are concise: all components have constant size, except \(|S|\) grows logarithmically with the number of ledger entries, and \(\mathcal{K}_r\) and \(\mathcal{K}_r\) have up to \(N-f\) entries. Moreover, most intermediate hashes in \(S\) can be shared across collections of receipts. We explored using signature aggregation [20] to reduce the size of \(\mathcal{K}_r\), but this is not worthwhile for realistic consortia sizes: verifying these signatures is significantly more expensive than our current implementation. Any third party can verify the receipt, as described in Alg. 2, which is similar to how the client verifies \(\text{reply}\) and \(\text{reply}\) messages.

### 3.3 Performance optimizations

L-PBFT supports checkpointing the server side. This is an important optimization: new replicas can start processing requests after obtaining a checkpoint without having to replay the ledger from the start; slow replicas can be brought up-to-
date using a recent checkpoint; and auditing can start from a checkpoint instead of the beginning of the ledger.

Checkpoints include the key-value store $kv$ and a compressed representation of the Merkle tree where all complete subtrees are replaced by their roots. Replicas start creating a checkpoint $cp_s$ when they execute a batch with sequence number $s$ such that $s \mod C = 0$. The primary adds a batch to the ledger at sequence number $s+C$ with a special checkpoint transaction that records the checkpoint digest. $C$ is chosen to give replicas enough time to complete a checkpoint without delaying the replication protocol. The backups only accept the pre-prepare for $s+C$ if they computed the same checkpoint digest for sequence number $s$.

When replicas fetch checkpoint $cp_s$, they also fetch the ledger up to $s$, but they do not need to replay the ledger or check all signatures (with the exception of governance transactions, see §5.2). They check the signatures in checkpoint receipts and that the contents of the ledger between consecutive checkpoints are consistent with the Merkle tree roots in the corresponding receipts. They do this from the start of the ledger until $s+C$. This allows auditing to assign blame accurately when correct replicas fetch malformed ledgers from misbehaving replicas with more than $f$ failures.

L-PBFT also uses techniques to reduce the cost of cryptography. All messages are sent over encrypted and authenticated connections. Using authenticated connections even for signed messages mitigates denial-of-service attacks that consume replica resources verifying signatures [28]. Signature verification is parallelized for both messages received from replicas and from clients [19, 28], which improves throughput and scalability significantly. To further optimize performance, L-PBFT overlaps execution of request batches with validating the pre-prepare signature. Backups decide whether to send the prepare only after both complete. Since the message is received over an authenticated connection from the primary, this always succeeds for correct primaries.

4 Auditing and enforcement

In this section, we describe how auditing produces universal proofs-of-misbehavior (uPoMs) when linearizability is violated (§4.1), and the role of the enforcer in obtaining ledgers for auditing and punishing the members responsible for misbehaving replicas (§4.2). We first focus on the simpler case of auditing without governance transactions; §5 describes governance transactions and their impact on auditing.

4.1 Auditing

Auditing is triggered when clients obtain a sequence of transaction receipts that violates linearizability, i.e., when no linearizable execution of the stored procedures that define the service transactions can produce the sequence of receipts. The set of receipts in such a sequence is said to be inconsistent. For example, Alice gets a receipt $\langle t, i, o, x \rangle$ where $t$ is “deposit a million dollars into Bob’s account” and $o$ is “debit successful”. Then Alice sends Bob the receipt and Bob runs a transaction to check his balance that returns a receipt $\langle t', i', o', x' \rangle$ where $t'$ is “get account balance” and $o'$ is “current balance is 100 dollars”. If $j \leq i$, linearizability is violated because Bob issued the balance check only after receiving the receipt from Alice. If $j > i$ and Bob did not withdraw at least 999,900 dollars from his account, linearizability is also violated. In either case, Bob can decide to start an audit to recover the lost funds.

Anyone can perform an audit given a set of receipts and the corresponding ledger segment. If the receipts are indeed inconsistent, the audit produces uPoMs for at least $f+1$ misbehaving replicas. The auditor can assign blame even if all replicas misbehave, e.g., they can all collude and rewrite the ledger but they cannot rewrite the receipts held by the clients. Alg. 3 describes the auditing process.

An auditor receives an ordered set of receipts $R = \langle \langle t_0, i_0, o_0, x_0 \rangle, \ldots, \langle t_k, i_k, o_k, x_k \rangle \rangle$ where $k \geq 1$ and $\forall l < k : i_l \leq i_{l+1}$. It invokes verifyReceipt (Alg. 2) to check if the receipts are valid (Alg. 3, lines 2–4). Then asks the enforcer to obtain the last checkpoint before $i_0$, which we denote by $C_0$ with sequence number $s_{C_0}$, and the ledger segment $L$ from $s_{C_0}$ up until the batch that commits $i_k$, all obtained from replica $r$ (line 5). If the ledger does not contain the batch that commits $i_k$ yet, the enforcer obtains the evidence that it committed from $M$ (Alg. 1) of the $N-f$ replicas that signed $x_k$.

After validating that $C_0$ matches the checkpoint digest in $x_0$, the auditor sets the service state to the checkpoint value (line 6) and starts replaying the ledger (line 7). Since PAC provides accountability even if all replicas fail, it is necessary to replay transaction execution to check if the results are correct. However, the auditor does not need to understand the semantics of the service; it can retrieve the code of the stored procedures from $C_0$.

Replaying the ledger is similar to running L-PBFT at a correct backup (Alg. 1, lines 16–25). It involves executing each batch of transactions, checking the commitment evidence, and checking that the root of the ledger Merkle tree matches the one in the pre-prepare message signed by the current primary (Alg. 1, line 8); view-change and new-view messages in the ledger allow the auditor to determine the primary at each

```
 Alg. 3: High-level auditing process

1. on audit(R) = \{\langle t_0, i_0, o_0, x_0 \rangle, \ldots, \langle t_k, i_k, o_k, x_k \rangle\}
2. foreach \( \langle i, t, o, x \rangle, x' \rangle \in R \) do
3. if not verifyReceipt(\( \langle i, t, o, x \rangle, x' \rangle \)) then
4. return invalidReceipt
5. \( C_0, x_{C_0}, L, r := \) getCheckpointAndLedger(x_{C_0}, x_0)
6. \( s_{cp} := x_{C_0}; cp := C_0; kv := loadCheckpoint(x_{C_0}, C_0) \)
7. foreach \( s \in x_{C_0}, \ldots, seqno(x) \) do
8. replayBatch(r, L, kv)
9. foreach \( \langle i, t, o, x, x' \rangle \in R \) seqno(x) = seqno(x')) do
10. if not verifyLedgerStateAndReceipt(\( L, kv, s \langle i, t, o, x, x' \rangle \)) then
11. \( F := createLedgerFragment(seqno(s, \langle s \rangle), \langle s \rangle) \)
12. uPoM := \( \langle i, F, cp, \langle i, t, o, x, x' \rangle \rangle \); return uPoM
13. if \( s \mod C = 0 \) then
14. \( s_{cp} := s; cp := createCheckpoint(kv) \)
15. return noMisbehavior
```
point. The auditor runs code similar to Alg. 1, lines 60–65, to check messages, and it also checks the correctness of checkpoints and their receipts. The variables $cp$ and $s_P$ track the last checkpoint replayed and its sequence number.

If replaying the ledger fails, it may not be replica $r$’s fault because $r$ may have fetched a checkpoint after the malformed ledger segment without having executed the segment. If the ledger segment matches the Merkle tree root in the receipt for the next checkpoint, the replicas that signed the checkpoint receipt are to blame; otherwise, only $r$ can be blamed, and the auditor engages the enforcer to ask another replica, which has not been blamed yet, for the rest of the ledger. In both cases, the auditor sends a uPoM $\langle i, F, cp, nil \rangle$ to the enforcer, where $i$ is the position at which the inconsistency was detected and $F$ is the ledger fragment between the last checkpoint before $i$ and the first checkpoint after $i$. This case is not shown in Alg. 3, and it is the only case in which the uPoM does not include a conflicting receipt.

The auditor also validates that the ledger replay is consistent with the transaction receipts (Alg. 3, line 10). If the sequence of receipts is inconsistent, as claimed, the ledger fails to match at least one of the receipts. For each receipt $\langle v, i, o_i, \rangle$, $\langle v, s, H(k_p), E_{i-P}, E_p, d_c, H(pp_p), \sigma_p, E_s, \Sigma_s, K_s, S \rangle$, the auditor checks that the Merkle root computed from the receipt (Alg. 2, lines 2–4) matches the Merkle root in the pre-prepare at sequence number $s$ in ledger $L$.

If the two roots do not match, there are two cases: (i) the pre-prepare with sequence number $s$ in $L$ has a view number $v_i \leq v$; or (ii) $v_i > v$. In case (i), the ledger contains evidence at sequence number $s + P$ that the batch with sequence number $s$ committed. Since at least $f + 1$ of the replicas that signed the evidence also signed the receipt, they can be blamed.

In case (ii), since $v < v_i$, there must be at least $N - f$ view-change messages from different replicas to view $v_i$ in $L$ that claim not to have prepared the batch in the receipt in prior views. Since there are at least $f + 1$ of those replicas that also signed the receipt, they can be blamed.

### 4.2 Enforcement

Since PAC provides accountability even if all replicas and members misbehave, we need an enforcer outside the system to obtain checkpoints and ledger segments for auditing, and to punish members responsible for misbehaving replicas. We expect that consortium members will sign a binding contract establishing penalties if a uPoM proves that a replica that they control misbehaved, or if they fail to produce checkpoints and ledgers for auditing within an agreed deadline. These penalties can be imposed by the enforcer, e.g., an arbitrator [13] or a court of law [14].

The auditor requests data for auditing from the enforcer by sending $x_0$ and $x_k$ (Alg. 3, line 5). The enforcer verifies that $x_0$ and $x_k$ are valid, and requests one of the replicas that signed $x_k$ for the last checkpoint before $x_0$, with sequence number $s_{C_0}$, and theledger segment from $s_{C_0}$ to the sequence number of the batch that commits $x_k$.

The replica’s checkpoint at $s_{C_0}$ may not match the checkpoint digest in $x_0$, but this does not imply that it is dishonest or that the replicas that signed $x_0$ are dishonest. To ensure that an honest replica is not blamed, the replica returns a receipt for the checkpoint at $s_{C_0}$ to the auditor. The auditor can blame the $N - 2f > f + 1$ replicas with signatures and nonces in both $x_0$ and the receipt for the different checkpoint.

Correct replicas have the information requested by the enforcer and can produce it within a short period. If the enforcer does not receive it from the replica within a reasonable time-out, it contacts the controlling consortium member to obtain the checkpoint and ledger. If the member fails to provide this information by an agreed deadline, it is punished according to the contract. This is important to ensure that misbehaving members cannot escape punishment by failing to produce information for auditing. However, it introduces a weak synchrony assumption that may lead to the punishment of honest members. We expect that the deadline will be set conservatively to make this unlikely in practice.

The enforcer also punishes members if a uPoM proves that a replica that they control misbehaved. When it receives a uPoM $\langle i, F, cp, \langle v, i, o_i, \rangle \rangle$, the enforcer checks its validity by replaying the ledger fragment $F$ from the checkpoint $cp$. It follows a procedure similar to Alg. 3, line 7-10, but the number of transactions to replay is bounded by the transactions between consecutive checkpoints. As an optimization, the number of transactions to replay is required only when the uPoM pertains to a ledger segment that is well-formed except for the incorrect execution of some transaction, which can happen only with more than $N - f$ misbehaving replicas. The uPoMs for all other cases can be verified more efficiently by replaying the protocol messages in the ledger fragment and reconstructing the Merkle tree at each point starting from the compressed Merkle tree in the checkpoint (without loading the full checkpoint of the key-value store).

If the uPoM is incorrect, the enforcer punishes the auditor; otherwise, it punishes the members responsible for at least $f + 1$ replicas. The enforcer also penalizes entities that request information for auditing and fail to produce a corresponding uPoM, or that produce a uPoM and fail to indicate that it can be verified using the optimization.

We expect auditing to be rare, because PAC provides linearizability with up to $f$ misbehaving replicas and because of the penalties on misbehaving auditors. Therefore, we expect the load placed on the enforcer to be small.

### 5 Governance

PAC extends CCF’s governance model [5] with support for accountability with any number of Byzantine faults. PAC maintains governance data that includes the public signing keys for members and replicas, an endorsement of each replica’s signing key signed by the member responsible, and the stored procedures to run transactions. We refer to the governance
data as the *constitution* of the service. To create a new instance of PAC, all members sign the initial constitution as part of a *genesis transaction*.

For an effective distribution of trust, the constitution should ensure replicas are evenly distributed across members. PAC provides the flexibility for a member to delegate replica management to other principals, but the member retains responsibility. The constitution should also prevent members from removing replicas that they are not responsible for.

### 5.1 Dynamic configuration

To change the constitution, consortium members use *governance transactions*, which have two types: propose and vote. Their behavior is defined by stored procedures, which can also be modified using governance transactions.

To make changes, a member invokes a propose transaction, which records a proposed change to the constitution. Other members then invoke the vote transaction to record their vote if the proposed changes should go ahead. Every time a vote is recorded the votes are recounted and if the proposed change has enough supporting votes, the transaction marks the proposal as approved and executes a stored procedure to modify the constitution.

Members can propose to add or remove replicas, which may cause the service to change. Replicas are identified by their public keys, and the configuration (i.e., the set of replicas running the service) is identified by a digest of the concatenation of the replicas’ public keys. If enough members vote to approve a proposal that changes the replica set, and the configuration did not change in the meantime, the final vote transaction changes the configuration. The signed proposal is used as the endorsement of a replica added by the proposing member.

When a primary executes an approved proposal at sequence number $s$, it ends the current batch. It then sends $P$ end-of-configuration *pre-prepare* messages with empty transaction batches. These do not affect the key-value store, but the *pre-prepare* for each sequence number $s'$ includes evidence that the batch for sequence number $s'$ committed ($\S$3). They are added to the ledger to prove that all transactions in batches up to and including $s$ have committed. This simplifies auditing by ensuring that all evidence is added to the ledger in the same configuration as the *pre-prepare*. The configuration change takes effect after sequence number $s+P$.

After the configuration change takes effect, the replicas in the new configuration create a checkpoint of the key-value store at sequence number $s+P$ and send the *pre-prepare* for the checkpoint transaction at $s+P+1$, followed by $P$ special start-of-configuration *pre-prepare* messages with empty transaction batches. This ensures that the evidence that the checkpoint transaction committed is recorded in the ledger before any other transactions are executed in the new configuration. Therefore, the checkpoint digests in *pre-prepare* messages and receipts always refer to checkpoints taken in the same configuration, i.e., with the same $N$ and signing keys. Changes to the stored procedures for transactions and consortium membership are similar.

New replicas must catch up by obtaining the ledger, a recent checkpoint, and replaying the ledger from that checkpoint ($\S$3.3). They start fetching the ledger and checkpoint before the proposal to reduce the impact on performance.

Replicas that are no longer part of the new configuration retire after receiving the *pre-prepare* for sequence number $s+P+1$. Members and replicas that are removed from PAC should delete their private signing keys to provide forward security. This prevents them from being blamed for future compromises while still allowing authentication of transactions in the ledger using their public keys.

### 5.2 Transaction receipts

PAC produces receipts for governance transactions and records them in the ledger, like other transactions. Since all governance transactions are recorded in the ledger, an auditor or replica can determine the constitution that is in effect at each point in the ledger. This allows the auditor to verify member and replica signatures and to determine what code to use to execute a transaction request. Clients, however, do not store a full copy of the ledger, but they need to verify signatures in receipts. PAC solves this problem by having clients store receipts for all governance transactions since the genesis transaction. Since governance transactions are rare, this is much smaller than the full ledger.

A client checks that a transaction receipt for index $i$ is valid by considering all governance transaction receipts from the genesis transaction up to $i$. It carries out an audit process, as described in $\S$4, but only for the governance transactions. If the audit succeeds, the client obtains the replica signing keys at index $i$, which it uses to validate the receipt ($\S$3).

This raises the challenge how a client determines that it has all the required governance transactions. PAC solves this problem by including the ledger index of the last governance transaction in every *pre-prepare* message and receipt. A client can thus request missing governance transaction receipts by traversing the sequence of governance transactions. Clients can cache governance transaction receipts and the result of auditing them, and they can fetch these receipts and audit them incrementally for improved performance.

### 5.3 Auditing

The auditing process in $\S$4.1 changes as follows. In addition to the sequence of possibly inconsistent client receipts, the auditor is also sent the supporting governance transaction receipts (with duplicates removed). Then *verifyReceipt* (Alg. 3, line 3) replays these governance transactions to determine $N$ and the signing keys to verify each client receipt. Most of this work can be reused across client receipts. If the auditor finds inconsistent governance transaction receipts, it can find proofs of misbehavior using Alg. 3 as for inconsistent client receipts. Verifying the consistency between the ledger and re-
receipts (Alg. 3, line 10) remains unchanged because of PAC’s implementation of configuration changes.

When the auditor requests data for auditing from an enforcer or sends a uPoM to an enforcer, it must also send the supporting governance transaction receipts.

6 Correctness

Next we give correctness arguments for L-PBFT and auditing.

**Theorem 6.1.** L-PBFT is linearizable.

**Proof (sketch).** L-PBFT extents the PBFT algorithm, which is linearizable [23]. L-PBFT adds early execution and nonce commitment. If these extensions do not affect linearizability, L-PBFT is also linearizable.

**Early execution.** Replicas execute transactions tentatively before a batch prepares. If a batch does not commit, L-PBFT rolls back the transactions. This is equivalent to PBFT, which executes transactions after the batch committed.

**Nonce commitment.** L-PBFT, like PBFT, signs pre-prepare and prepare messages; unlike PBFT, L-PBFT does not sign commit messages. Replicas sample a fresh random nonce for each pre-prepare or prepare message with sequence number $s$ at view $v$, and add a hash of this nonce to the signed payloads. Later in the protocol, replicas include the nonce in the commit message, instead of an extra signature. We show that this provides the same standard cryptographic security as the signature scheme (namely, resistance to existential forgery against chosen-message attacks) as long as the cryptographic hash function is second pre-image resistant on random inputs. Since the addition of a nonce to the signed payloads is injective, a forgery of a L-PBFT authenticator for a pre-prepare or prepare message yields a forgery against the signature scheme. A forgery of an authenticator for a commit message, i.e., a value with the same hash as a fresh random nonce that has not yet been revealed, is a second pre-image collision.

**Theorem 6.2.** Auditing can assign blame to at least $f+1$ misbehaving replicas if a true linearizability violation is reported.

**Proof (sketch).** Assume the auditor obtains a sequence of valid client receipts (with backing governance transaction receipts linking then to the genesis transaction) that violate linearizability. Let $s_{C_0}$ be the sequence number of the checkpoint with digest $d_{C_0}$ in the first receipt and $s_k$ be the sequence number of the last receipt.

The auditor, via the enforcer, asks the replicas in $R_k$ that signed the last receipt or the last governance transaction (whichever comes later) for the checkpoint at $s_{C_0}$, and the ledger entries from $s_{C_0}$ to $s_k+P$ (or $s_k$ plus evidence from $M$ in Alg. 1 if the ledger does not extend to $s_k+P$). Since any correct replica in $R_k$ is guaranteed to have this information, the enforcer assigns blame to replicas that fail to provide it.

A correct replica may have a checkpoint at $s_{C_0}$ with a digest different from $d_{C_0}$. Such a replica returns the checkpoint receipt, and the enforcer can assign blame to the $f+1$ replicas that signed the checkpoint receipt and the first receipt. A correct replica may also have fetched a malformed ledger fragment from misbehaving replicas. In this case, the auditor assigns blame to the $N-f$ replicas that signed the checkpoint receipt at the end of the malformed fragment.

Therefore, the auditor either obtains a checkpoint at $s_{C_0}$ with digest $d_{C_0}$ and a well-formed ledger, or blames at least $f+1$ replicas. With a valid checkpoint and ledger, the auditor can assign blame to at least $f+1$ replicas when a receipt is inconsistent with the ledger entry at the same sequence number (see §4.1). If the receipt sequence violates serializability, there is always at least one inconsistency because the ledger records a valid serial history; otherwise, a linearizability violation implies that a client received a receipt proving execution at a position $j \leq i$ for a request that it sent after receiving the receipt for $i$. Since the receipt for $i$ includes a signature of ledger’s Merkle tree root and requests cannot be forged, there will be at least one inconsistency at $j \neq i$.

7 Evaluation

We evaluate PAC to understand the cost of providing receipts (§7.1), its scalability (§7.2), the overheads of receipt validation (§7.3), and auditing (§7.4). We finish with a breakdown of the impact of PAC’s design features (§7.5).

**Testbeds.** Our experimental setup consists of three environments: (a) a dedicated cluster with 16 machines, each with an 8-core 3.7-GHz Intel E-2288G CPU with 16 GB of RAM and a 40 Gbps network with full bi-section bandwidth; (b) a LAN environment in the Azure cloud, with Fsv2-series VMs with 16-core 2.7-GHz Intel Xeon 8168 CPUs and 7 Gbps network links; and (c) a WAN environment with the same VMs across 3 Azure regions (US East, US West 2, US South Central). All machines run Ubuntu Linux 18.04.4 LTS.

**Implementation.** Our PAC implementation has approx. 40,000 lines of C++ code. It uses the formally-verified Merkle trees and SHA functions of EverCrypt [52], the MbedTLS library [46] for client connections, secp256k1 [60] for all secure signatures, and replicas create secure communication channels via a Diffie–Hellman key exchange.

**Benchmarks.** We use the SmallBank benchmark [9], which models a banking application in which customers deposit, withdraw, and transfer funds across 1,000,000 accounts. Clients randomly execute 5 transaction types that deposit and withdraw funds, check account balances, and amalgamate accounts. Each request is between 300–400 bytes.

Transaction throughput is measured at the primary replica and latency at the clients. All experiments are compute-bound. Results are averaged over 5 runs, with min/max error bars. Since PAC uses well-established BFT techniques with fewer than $N/3$ failures, we only report results for failure-free execution and focus on understanding the performance impact of producing receipts to assign blame.
Fig. 4: Transaction throughput/latency ($f=1$, dedicated cluster)

**Baselines.** We compare against four baselines: (i) PAC-PeerReview, which uses PeerReview’s mechanism for blame assignment [34]. Replicas sign all messages and send signed acknowledgements on message receipt; (ii) PAC-NoReceipt, a PAC variant that produces a ledger but no receipts; (iii) HotStuff [61], a state-of-the-art BFT protocol, which is at the core of the Diem permissioned ledger system [10]; and (iv) Hyperledger Fabric (v. 2.2) [11], the most widely deployed open-source permissioned ledger system. Fabric’s latest major release does not include a BFT consensus protocol [40] but only supports crash-stop failures with Raft [51]. We compare against this computationally less expensive version.

### 7.1 Transaction throughput and latency

We explore the throughput and latency of transaction execution with 4 replicas ($f=1$) in the dedicated cluster, comparing PAC, PAC-NoReceipt, PAC-PeerReview, and Fabric.

Fig. 4 shows the throughput/latency plot as more transaction requests are sent. PAC achieves 48,034 tx/s while maintaining latencies below 40 ms. As the load increases, it becomes saturated with queuing delays adding to the latency. PAC-NoReceipt’s peak throughput is 49,516 tx/s, which is only 3% higher than PAC, demonstrating the low overhead of receipt generation.

PAC-PeerReview exhibits an order of magnitude lower throughput, because all messages are signed and include the hash of a per-replica ledger of all messages. Signatures cannot be re-used when the same payload is sent to multiple recipients, e.g., a replica must sign a reply message for every transaction in a batch. PAC-PeerReview therefore performs two orders of magnitude more asymmetric cryptographic operations than PAC.

Fabric’s throughput is 1,207 tx/s, with a latency of 1.6 s, which is substantially worse than PAC, despite not using a BFT protocol. This shows that for full permissioned ledger systems often overheads such as running transactions against a key-value store dominate [50].

### 7.2 Scalability

Next we consider the effect on PAC’s transaction throughput when increasing the number of replicas (and tolerated failures $f$) in the Azure WAN environment, which spans multiple regions to reduce correlated failures [16].

Fig. 5 shows that PAC’s throughput decreases with more replicas because more signatures must be verified by each replica. Since each replica has a fixed number of threads that can validate signed pre-prepare/prepare messages in parallel, throughput decreases when the replica count exceeds the number of hardware threads (16). PAC, however, is only marginally affected by the higher WAN latencies due to its use of pipelining, as shown by the comparison to PAC deployed in the Azure LAN environment.

In contrast, HotStuff [3] achieves only a peak throughput of 5,862 tx/s in the WAN environment, which is significantly worse than its LAN throughput [4]. While it degrades slowly with more replicas, even with 64 replicas its throughput remains 71% lower than that of PAC. The throughput of PAC-PeerReview is even lower because it is bottlenecked by its more extensive cryptographic operations.

We want to study the difference in WAN performance between PAC and HotStuff in more detail. For this, we measure the request latency of both systems under low load, which ensures that latency is not dominated by queuing time at replicas. Table 1 shows that HotStuff’s request latency is twice that of PAC’s. For both systems, request latency is dominated by the required number of network round trips. This shows the benefit of clients in PAC receiving transaction results with receipts in only 2 round trips.

### 7.3 Receipt validation

The size of receipts and their computation overhead are dominated by two factors: the size of the Merkle tree when a receipt is created and the number of replicas that sign it. We consider two scenarios in which PAC runs for 1 and 100 years, respectively, while executing 50,000 tx/s.

After 1 year, the evidence for the Merkle tree path is 1,280 bytes and, on average, can be calculated in 9 µs; after 100 years, it is 1,472 bytes and can be calculated in 10 µs. We observe that, in both cases, the cost is dominated by the signature verification, which takes 18 ms and 52 ms for $f=1$ and $f=3$, respectively.

### 7.4 Ledger auditing

Next we want to understand the performance of auditing a ledger. We use SmallBank to compare PAC’s execution to
ledger auditing. When measuring throughput at $f=1$, auditing is 23% faster than execution, because there is no network overhead, message signing, or ledger writes. In each batch, we only verify $2f+1$ rather than up to $3f+1$ signatures. For $f=4$, the performance gap increases to 67%, as more replicas add to the communication and cryptographic load during execution. Higher values of $f$ present additional opportunities for parallelizing cryptographic verification.

### 7.5 Breakdown of PAC features

To provide receipts, PAC requires cryptographic operations that go beyond the ones in traditional BFT state machine replication protocols: it produces a ledger, maintains a Merkle tree, generates transaction receipts, etc. Here we explore the impact of these features on PAC’s throughput in the dedicated cluster.

We compare several variants of PAC, each limiting functionality further: (a) full PAC; (b) PAC-NoReceipt, i.e., without creating or sending receipts; (c) without signed client requests, i.e., transactions in the ledger cannot be linked to clients; (d) using MACs for message authentication between replicas; (e) without a ledger; and (f) with empty requests only, i.e., without the overhead of executing transactions against the key-value store (equivalent to plain PBFT).

| Variant                        | Throughput (tx/s) |
|--------------------------------|-------------------|
| (a) Full PAC                   | 48,034            |
| (b) PAC-NoReceipt              | 49,516            |
| (c) + without signed client requests | 108,827       |
| (d) + with MACs only           | 128,117           |
| (e) + without ledger           | 133,106           |
| (f) + with empty requests      | 299,227           |
| HotStuff (with empty requests) | 307,997           |
| Pompê (with empty requests)    | 465,646           |

Table 2 shows that full PAC (a) and PAC-NoReceipt (b) have a comparable throughput of 48,034 tx/s and 49,516 tx/s, respectively, but not verifying client signatures (c) doubles the throughput. Only using MACs instead of signatures (d) or removing the ledger altogether (e) do not increase throughput substantially. Removing the overhead of request execution against the key-value store (f), doubles the throughput again.

For context, we compare to two Byzantine consensus protocols with similar guarantees to (f) above, HotStuff [61] and Pompê [4, 63]. HotStuff’s throughput is 307,997 tx/s, but it with higher latency ($\S$7.2). By separating request ordering and consensus, Pompê achieves a throughput of 465,646 tx/s, also with worse latency (PAC’s 12 ms to Pompê’s 73 ms). PAC could utilize Pompê’s techniques for increased throughput by sacrificing its two round-trip latency.

These breakdown results show that PAC’s overhead comes primarily from the cryptographic operations required for verifying client requests, and the transactional key-value store, rather than the consensus protocol or the mechanisms to assign blame.

### 8 Related work

**Permissioned ledgers.** A number of permissioned ledger systems [10, 11, 39, 53] rely on BFT consensus protocols to agree on an order for transactions. Hyperledger Fabric [40], HyperLedger Besu [39] and Quorum [53] use variants of PBFT [49, 56], which do not retain proof of a replica’s operations, and therefore cannot assign blame. Diem [10] uses the DiemBft [17] consensus protocol, which is based on HotStuff [61] and lacks accountability features.

**Byzantine consensus** [24, 28, 43] distributes trust. Recent work on BFT protocols has focused on improving guarantees [12, 29, 48] or performance for particular use cases [57, 63]. SBFT [32] and HotStuff [61] scale to hundreds of replicas using threshold cryptography, which prevents blame assignment. For permissioned ledgers, scaling to many replicas without growing the consortium does not improve trustworthiness, and consortia typically cannot grow arbitrarily.

Other work has explored misbehavior and its impact on Byzantine consensus. BFT2F [44] formalizes safety and liveness guarantees after more than $f$ replicas are compromised. It provides PBFT’s guarantees with up to $f$ failures and provides fork* consistency with up to $2f$ failures. For permissioned ledgers, fork* consistency is not sufficient because it is susceptible to double-spending attacks.

Depot [45] issues proofs-of-misbehavior after observing malicious actions, but it adopts a form of eventual consistency, which is incompatible with permissioned ledgers. Pompê [63] prevents dishonest primaries from controlling the ordering of requests. It does not address scenarios in which there are more than $f$ dishonest replicas.

**Accountability.** PeerReview [34] ensures that nodes in a distributed system remain accountable for their actions. As shown in $\S$7.1, due to its generality, PeerReview incurs a high overhead when applied to a permissioned ledger. In contrast, PAC introduces mechanisms specific to BFT state machine replication, such as a shared ledger with a Merkle tree, to improve both regular transaction execution and auditing.

Accountable virtual machines [33] carries out auditing through spot checking of checkpoints, but has the same performance overheads as PeerReview. SNP [64] is a networking-specific implementation of accountability, offering provenance for routing decisions. Such specializations of PeerReview improve performance in particular domains, but are not directly applicable to permissioned ledgers.

BAR [8] incentivizes replicas to act honestly: honest replicas impose disincentives on misbehavior. This weaker model allows Bar to tolerate more than $1/3$ of replica failures. If these incentives fail [38], however, replicas share the blame.

Accountability with more than $f+1$ misbehaving replicas has been discussed previously [21, 22, 35]. BFT Protocol Forensics [41] and Polygraph [26] propose a ledger auditing mechanism, but assume that fewer than $N-f$ replicas misbe-
have and do not support changing replica sets. ASMR [54] offers accountability if fewer than \(N-f\) replicas misbehave, and limits the possible candidates when changing its replica set to a predetermined list.

9 Conclusions

In permissioned ledger systems, accountability is a strong disincentive for misbehavior. PAC is a permissioned ledger system that provides the evidence required to prove that \(f+1\) or more replicas misbehaved when clients observe safety violations (even if all replicas fail). It offers strong consistency and security properties while providing start-of-the-art performance compared to existing ledgers with weaker security guarantees. PAC achieves this by integrating collection of evidence to prove blame with a novel ledger-based BFT consensus algorithm.

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