Principles of Distributed Systems

inft-3507

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Lectures are based on: https://www.distributed-systems.net/index.php/books/ds4/

Chapter 08: Fault Tolerance

Dependability

Basics

A component provides services to clients. To provide services, the component may require the services from other components $\Rightarrow$ a component may depend on some other component.

Specifically

A component C depends on C^* if the correctness of C's behavior depends on the correctness of C^*'s behavior. (Components are processes or channels.)

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Requirements related to dependability

Requirement	Description
Availability	Readiness for usage
Reliability	Continuity of service delivery
Safety	Very low probability of catastrophes
Maintainability	How easy can a failed system be repaired

Reliability versus availability

Reliability $R(t)$ of component C

Conditional probability that C has been functioning correctly during $[0,t)$ given C was functioning correctly at time $T=0$.

Traditional metrics

- Mean Time To Failure (*MTTF*): The average time until a component fails.
- Mean Time To Repair (*MTTR*): The average time needed to repair a component.
- Mean Time Between Failures (*MTBF*): Simply $MTTF + MTTR$.

Reliability versus availability

Availability $A(t)$ of component C

Average fraction of time that C has been up-and-running in interval $[0,t)$.

- Long-term availability A: $A(\infty)$
- Note: $A = \frac{MTTF}{MTBF} = \frac{MTTF}{MTTF+MTTR}$

Observation

Reliability and availability make sense only if we have an accurate notion of what a failure actually is.

Terminology

Failure, error, fault

Term	Description	Example
Failure	A component is not living up to its specifications	Crashed program
Error	Part of a component that can lead to a failure	Programming bug
Fault	Cause of an error	Sloppy programmer

Terminology

Handling faults

Term	Description	Example
Fault prevention	Prevent the occurrence of a fault	Don't hire sloppy programmers
Fault tolerance	Build a component such that it can mask the occurrence of a fault	Build each component by two independent programmers
Fault removal	Reduce the presence, number, or seriousness of a fault	Get rid of sloppy programmers
Fault forecasting	Estimate current presence, future incidence, and consequences of faults	Estimate how a recruiter is doing when it comes to hiring sloppy programmers

Failure models

Types of failures

Type	**Description of server's behavior**
Crash failure	Halts, but is working correctly until it halts
Omission failure	Fails to respond to incoming requests
Receive omission	Fails to receive incoming messages
Send omission	Fails to send messages
Timing failure	Response lies outside a specified time interval
Response failure	Response is incorrect
Value failure	The value of the response is wrong
State-transition failure	Deviates from the correct flow of control
Arbitrary failure	May produce arbitrary responses at arbitrary times

Dependability versus security

Omission versus commission

Arbitrary failures are sometimes qualified as malicious. It is better to make the following distinction:

- Omission failures: a component fails to take an action that it should have taken
- Commission failure: a component takes an action that it should not have taken

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Observation

Note that deliberate failures, be they omission or commission failures, are typically security problems. Distinguishing between deliberate failures and unintentional ones is, in general, impossible.

Halting failures

Scenario

C no longer perceives any activity from C^* — a halting failure? Distinguishing between a crash or omission/timing failure may be impossible.

Asynchronous versus synchronous systems

- Asynchronous system: no assumptions about process execution speeds or message delivery times $\rightarrow$ cannot reliably detect crash failures.
- Synchronous system: process execution speeds and message delivery times are bounded $\rightarrow$ we can reliably detect omission and timing failures.
- In practice we have partially synchronous systems: most of the time, we can assume the system to be synchronous, yet there is no bound on the time that a system is asynchronous $\rightarrow$ can normally reliably detect crash failures.

Halting failures

Assumptions we can make

Halting type	Description
Fail-stop	Crash failures, but reliably detectable
Fail-noisy	Crash failures, eventually reliably detectable
Fail-silent	Omission or crash failures: clients cannot tell what went wrong
Fail-safe	Arbitrary, yet benign failures (i.e., they cannot do any harm)
Fail-arbitrary	Arbitrary, with malicious failures

Redundancy for failure masking

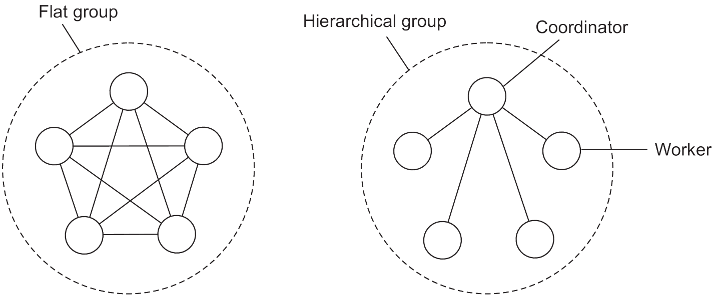
Types of redundancy

- Information redundancy: Add extra bits to data units so that errors can recovered when bits are garbled.
- Time redundancy: Design a system such that an action can be performed again if anything went wrong. Typically used when faults are transient or intermittent.
- Physical redundancy: add equipment or processes in order to allow one or more components to fail. This type is extensively used in distributed systems.

Process resilience

Basic idea

Protect against malfunctioning processes through process replication, organizing multiple processes into a process group. Distinguish between flat groups and hierarchical groups.



Groups and failure masking

k-fault tolerant group

When a group can mask any k concurrent member failures (k is called degree of fault tolerance).

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How large does a k-fault tolerant group need to be?

- With halting failures (crash/omission/timing failures): we need a total of $k+1$ members as no member will produce an incorrect result, so the result of one member is good enough.
- With arbitrary failures: we need $2k+1$ members so that the correct result can be obtained through a majority vote.

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- With arbitrary failures: we need $2k+1$ members so that the correct result can be obtained through a majority vote.

Important assumptions

- All members are identical
- All members process commands in the same order

Result: We can now be sure that all processes do exactly the same thing.

Consensus

Prerequisite

In a fault-tolerant process group, each nonfaulty process executes the same commands, and in the same order, as every other nonfaulty process.

Reformulation

Nonfaulty group members need to reach consensus on which command to execute next.

Flooding-based consensus

System model

- A process group $\mathbf{P} = \{P_1, \ldots, P_n\}$
- Fail-stop failure semantics, i.e., with reliable failure detection
- A client contacts a P_i requesting it to execute a command
- Every P_i maintains a list of proposed commands

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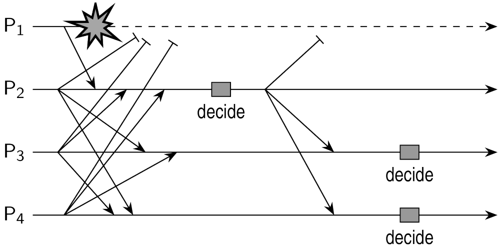
Basic algorithm (based on rounds)

1. In round r, P_i multicasts its known set of commands $\mathbf{C_i^r}$ to all others
2. At the end of r, each P_i merges all received commands into a new $\mathbf{C_i^{r+1}}$.
3. Next command cmd_i selected through a globally shared, deterministic function: $cmd_i \leftarrow select(\mathbf{C_i^{r+1}})$.

Flooding-based consensus: Example



Observations

- P_2 received all proposed commands from all other processes $\Rightarrow$ makes decision.
- P_3 may have detected that P_1 crashed, but does not know if P_2 received anything, i.e., P_3 cannot know if it has the same information as P_2 $\Rightarrow$ cannot make decision (same for P_4).

Raft

Developed for understandability

- Uses a fairly straightforward leader-election algorithm (see Chp. 5). The current leader operates during the current term.
- Every server (typically, five) keeps a log of operations, some of which have been committed. A backup will not vote for a new leader if its own log is more up to date.
- All committed operations have the same position in the log of each respective server.
- The leader decides which pending operation is to be committed next $\Rightarrow$ a primary-backup approach.

Raft

When submitting an operation

- A client submits a request for operation o.
- The leader appends the request $\langle o, t, \rangle$ to its own log (registering the current term t and length of).
- The log is (conceptually) broadcast to the other servers.
- The others (conceptually) copy the log and acknowledge the receipt.
- When a majority of acks arrives, the leader commits o.

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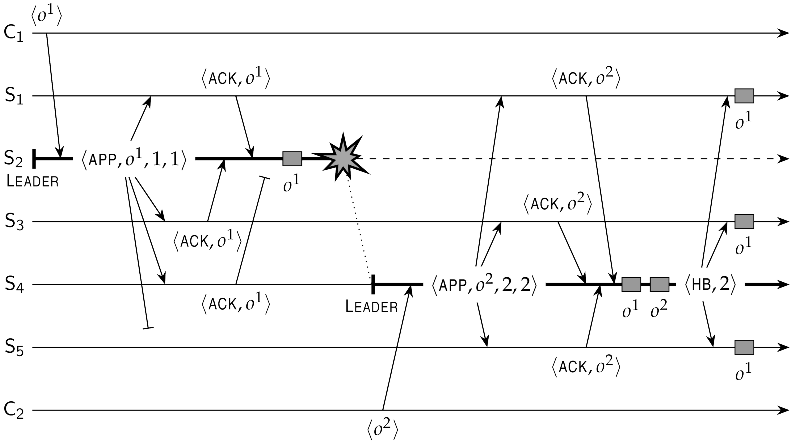
Note

In practice, only updates are broadcast. At the end, every server has the same view and knows about the c committed operations. Note that effectively, any information at the backups is overwritten.

Raft: when a leader crashes



Crucial observations

- The new leader has the most committed operations in its log.
- Any missing commits will eventually be sent to the other backups.

Realistic consensus: Paxos

Assumptions (rather weak ones, and realistic)

- A partially synchronous system (in fact, it may even be asynchronous).
- Communication between processes may be unreliable: messages may be lost, duplicated, or reordered.
- Corrupted message can be detected (and thus subsequently ignored).
- All operations are deterministic: once an execution is started, it is known exactly what it will do.
- Processes may exhibit crash failures, but not arbitrary failures.
- Processes do not collude.

Understanding Paxos

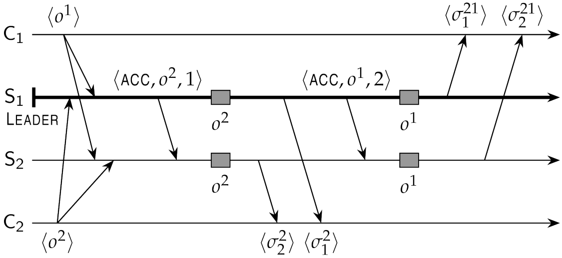
We will build up Paxos from scratch to understand where many consensus algorithms actually come from.

Paxos essentials

Starting point

- We assume a client-server configuration, with initially one primary server.
- To make the server more robust, we start with adding a backup server.
- To ensure that all commands are executed in the same order at both servers, the primary assigns unique sequence numbers to all commands. In Paxos, the primary is called the leader.
- Assume that actual commands can always be restored (either from clients or servers) $\Rightarrow$ we consider only control messages.

Two-server situation

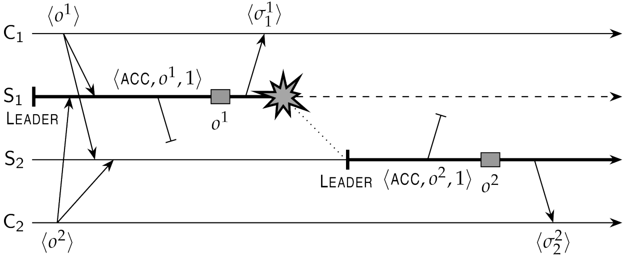


Handling lost messages

Some Paxos terminology

- The leader sends an accept message $\text{ACCEPT}(o, t)$ to backups when assigning a timestamp t to command o.
- A backup responds by sending a learn message: $\text{LEARN}(o, t)$
- When the leader notices that operation o has not yet been learned, it retransmits $\text{ACCEPT}(o, t)$ with the original timestamp.

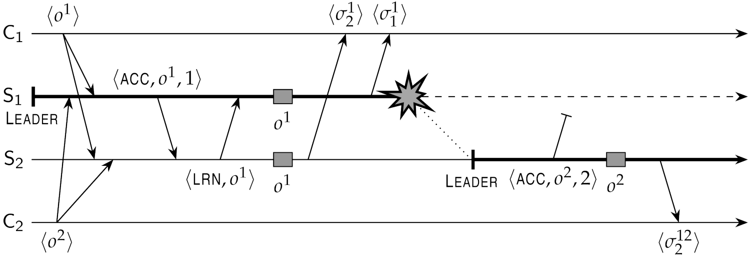
Two servers and one crash: problem



Problem

Primary crashes after executing an operation, but the backup never received the accept message.

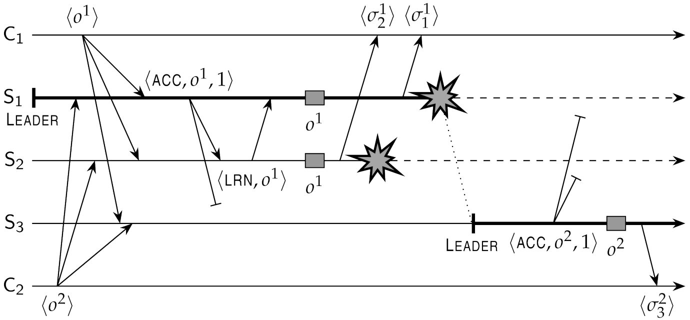
Two servers and one crash: solution



Solution

Never execute an operation before it is clear that is has been learned.

Three servers and two crashes: still a problem?



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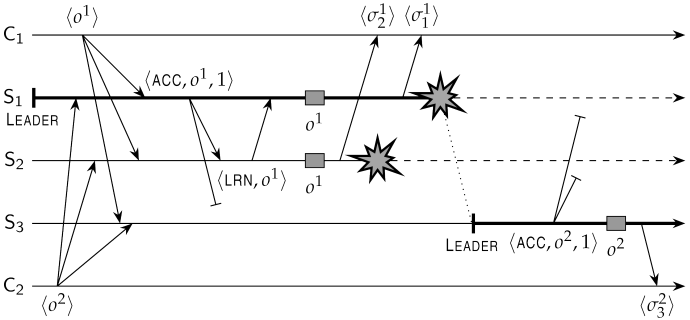

Scenario

What happens when $\text{LEARN}(o^1)$ as sent by S_2 to S_1 is lost?

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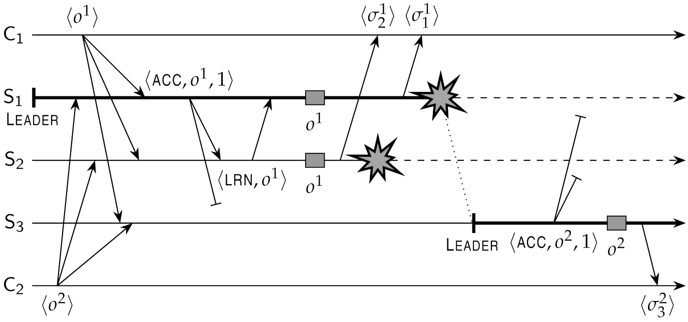

Scenario

What happens when LEARN(o^1) as sent by S_2 to S_1 is lost?

Solution

S_2 will also have to wait until it knows that S_3 has learned o^1.

Paxos: fundamental rule

General rule

In Paxos, a server S cannot execute an operation o until it has received a LEARN(o) from all other nonfaulty servers.

Failure detection

Practice

Reliable failure detection is practically impossible. A solution is to set timeouts, but take into account that a detected failure may be false.

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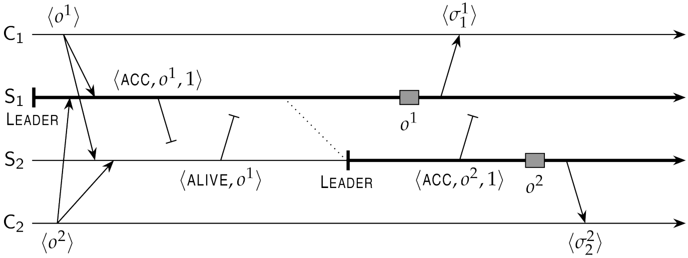

Required number of servers

Observation

Paxos needs at least three servers

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Adapted fundamental rule

In Paxos with three servers, a server S cannot execute an operation o until it has received at least one (other) LEARN(o) message, so that it knows that a majority of servers will execute o.

Required number of servers

Assumptions before taking the next steps

- Initially, S_1 is the leader.
- A server can reliably detect it has missed a message, and recover from that miss.
- When a new leader needs to be elected, the remaining servers follow a strictly deterministic algorithm, such as $S_1 \rightarrow S_2 \rightarrow S_3$.
- A client cannot be asked to help the servers to resolve a situation.

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Observation

If either one of the backups (S_2 or S_3) crashes, Paxos will behave correctly: operations at nonfaulty servers are executed in the same order.

Leader crashes after executing o^1

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S_3 is completely ignorant of any activity by S_1

- S_2 received $\text{ACCEPT}(o, 1)$, detects crash, and becomes leader.
- S_3 even never received $\text{ACCEPT}(o, 1)$.
- If S_2 sends $\text{ACCEPT}(o^2, 2) \Rightarrow S_3$ sees unexpected timestamp and tells S_2 that it missed o^1.
- S_2 retransmits $\text{ACCEPT}(o^1, 1)$, allowing S_3 to catch up.

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S_2 missed $\text{ACCEPT}(o^1,1)$

- S_2 did detect crash and became new leader
- If S_2 sends $\text{ACCEPT}(o^1,1) \Rightarrow S_3$ retransmits $\text{LEARN}(o^1)$.
- If S_2 sends $\text{ACCEPT}(o^2,1) \Rightarrow S_3$ tells S_2 that it apparently missed $\text{ACCEPT}(o^1,1)$ from S_1, so that S_2 can catch up.

Leader crashes after sending $\text{ACCEPT}(o^1, 1)$

S_3 is completely ignorant of any activity by S_1

As soon as S_2 announces that o^2 is to be accepted, S_3 will notice that it missed an operation and can ask S_2 to help recover.

S_2 had missed $\text{ACCEPT}(o^1, 1)$

As soon as S_2 proposes an operation, it will be using a stale timestamp, allowing S_3 to tell S_2 that it missed operation o^1.

Leader crashes after sending ACCEPT$(o^1, 1)$

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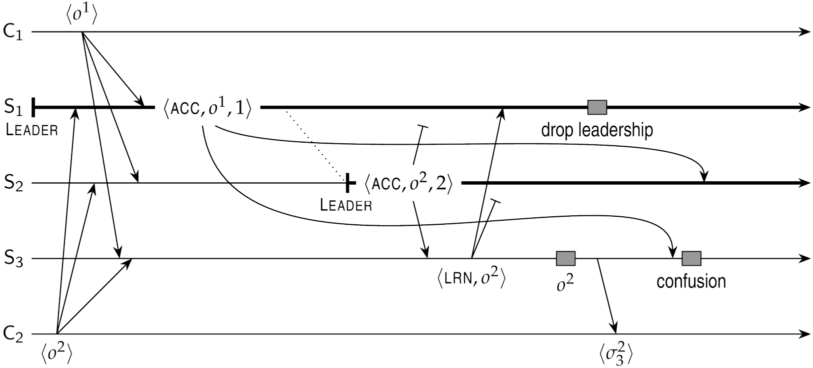
Observation

Paxos (with three servers) behaves correctly when a single server crashes, regardless when that crash took place.

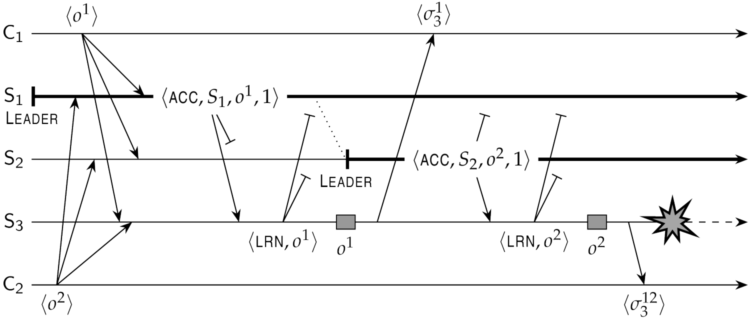
False crash detections



Problem and solution

S_3 receives ACCEPT$(o^1, 1)$, but much later than ACCEPT$(o^2, 1)$. If it knew who the current leader was, it could safely reject the delayed accept message $\Rightarrow$ leaders should include their ID in messages.

But what about progress?



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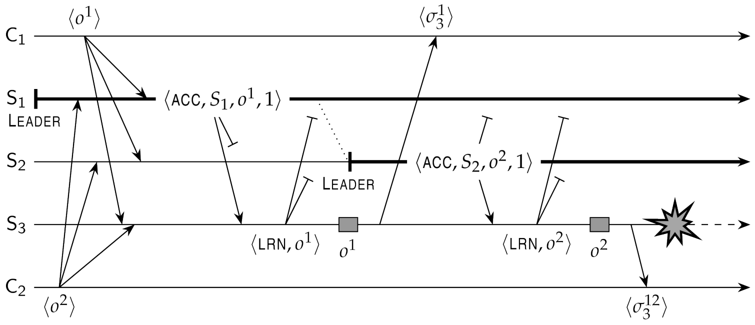


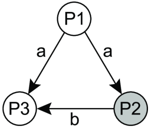
Essence of solution

When S_2 takes over, it needs to make sure that any outstanding operations initiated by S_1 have been properly flushed, i.e., executed by enough servers. This requires an explicit leadership takeover by which other servers are informed before sending out new accept messages.

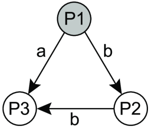
Consensus under arbitrary failure semantics

Essence

We consider process groups in which communication between process is inconsistent.



Improper forwarding



Different messages

Consensus under arbitrary failure semantics

System model

- We consider a primary P and $n-1$ backups $B_1, \ldots, B_{n-1}$.
- A client sends $v \in \{T, F\}$ to P
- Messages may be lost, but this can be detected.
- Messages cannot be corrupted beyond detection.
- A receiver of a message can reliably detect its sender.

Byzantine agreement: requirements

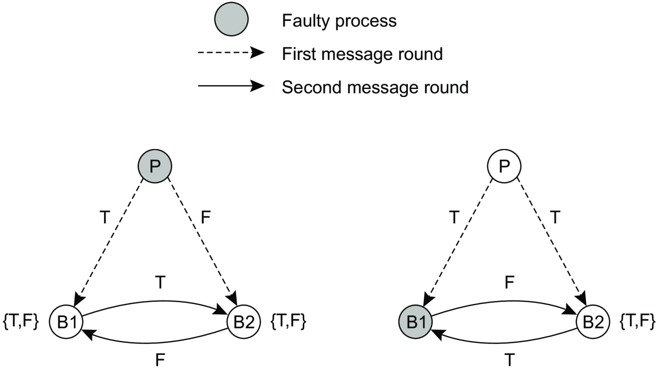
BA1: Every nonfaulty backup process stores the same value.

BA2: If the primary is nonfaulty then every nonfaulty backup process stores exactly what the primary had sent.

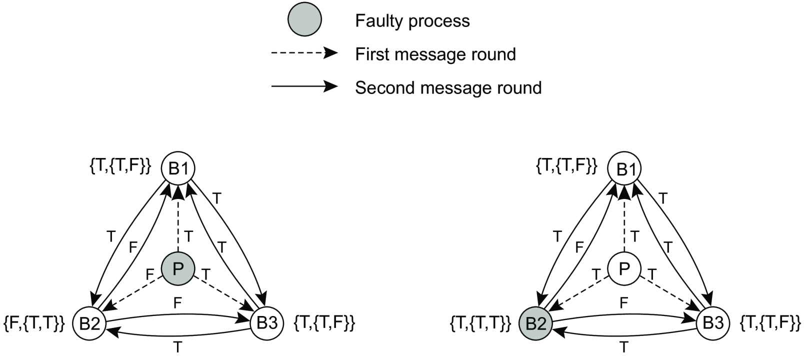
Observation

- Primary faulty $\Rightarrow$ BA1 says that backups may store the same, but different (and thus wrong) value than originally sent by the client.
- Primary not faulty $\Rightarrow$ satisfying BA2 implies that BA1 is satisfied.

Why having **3k** processes is not enough



Why having $3k+1$ processes is enough



Practical Byzantine Fault Tolerance (PBFT)

Background

One of the first solutions that managed to Byzantine fault tolerance while keeping performance acceptable. Popularity has increased with the introduction of permissioned blockchains.

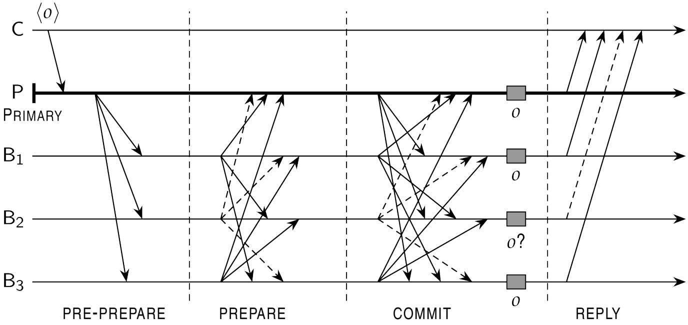
Assumptions

- A server may exhibit arbitrary failures
- Messages may be lost, delayed, and received out of order
- Messages have an identifiable sender (i.e., they are signed)
- Partially synchronous execution model

Essence

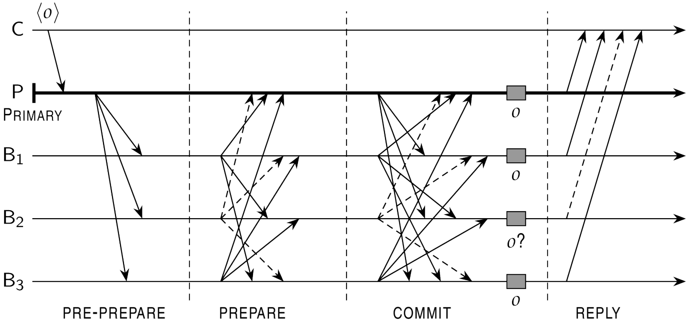
A primary-backup approach with $3k+1$ replica servers.

PBFT: four phases



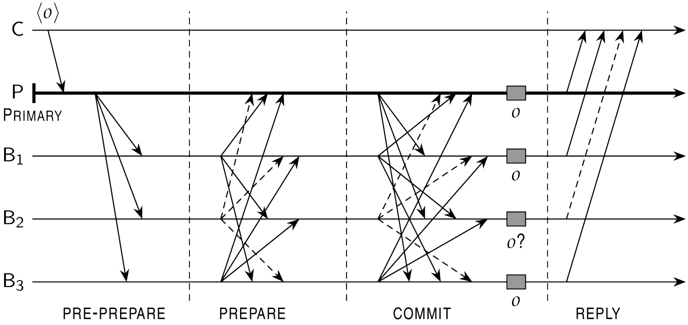
- C is the client
- P is the primary
- B_1, B_2, B_3 are backups
- Assume B_2 is faulty

PBFT: four phases



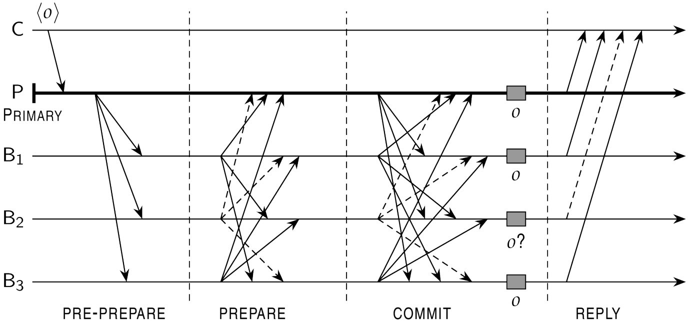
- All servers assume to be working in a current view v.
- C requests operation o to be executed
- P timestamps o and sends PRE-PREPARE(t, v, o)
- Backup B_i accepts the pre-prepare message if it is also is in v and has not accepted a an operation with timestamp t before.

PBFT: four phases



- B_i broadcasts $\text{PREPARE}(t, v, o)$ to all (including the primary)
- Note: a nonfaulty server will eventually log $2k$ messages $\text{PREPARE}(t, v, o)$ (including its own) $\Rightarrow$ consensus on the ordering of o.
- Note: it doesn't matter what faulty B_2 sends, it cannot affect joint decisions by P, B_1, B_3.

PBFT: four phases



- All servers broadcast COMMIT(t, v, o)
- The commit is needed to also make sure that o can be executed now, that is, in the current view v.
- When $2k$ messages have been collected, excluding its own, the server can safely execute o en reply to the client.

PBFT: when the primary fails

Issue

When a backup detects the primary failed, it will broadcast a view change to view $v+1$. We need to ensure that any outstanding request is executed once and only once by all nonfaulty servers. The operation needs to be handed over to the new view.

Procedure

- The next primary P^* is known deterministically
- A backup server broadcasts VIEW-CHANGE$(v+1,\mathbf{P})$: $\mathbf{P}$ is the set of prepares it had sent out.
- P^* waits for $2k+1$ view-change messages, with $\mathbf{X} = \bigcup \mathbf{P}$ containing all previously sent prepares.
- P^* sends out NEW-VIEW$(\mathrm{v}+1,\mathbf{X},\mathbf{O})$ with $\mathbf{O}$ a new set of pre-prepare messages.
- Essence: this allows the nonfaulty backups to replay what has gone on in the previous view, if necessary, and bring o into the new view $v+\mathit{1}$.

Realizing fault tolerance

Observation

Considering that the members in a fault-tolerant process group are so tightly coupled, we may bump into considerable performance problems, but perhaps even situations in which realizing fault tolerance is impossible.

Question

Are there limitations to what can be readily achieved?

- What is needed to enable reaching consensus?
- What happens when groups are partitioned?

Distributed consensus: when can it be reached

Process behavior	Commun. delay	Message ordering: Unordered — Message transmission: Unicast	Unordered — Multicast	Ordered — Unicast	Ordered — Multicast
Synchronous	Bounded	✓	✓	✓	✓
	Unbounded			✓	✓
Asynchronous	Bounded				✓
	UnBounded				✓

Formal requirements for consensus

- Processes produce the same output value
- Every output value must be valid
- Every process must eventually provide output

Consistency, availability, and partitioning

CAP theorem

Any networked system providing shared data can provide only two of the following three properties:

- C: consistency, by which a shared and replicated data item appears as a single, up-to-date copy
- A: availability, by which updates will always be eventually executed
- P: Tolerant to the partitioning of process group.

Conclusion

In a network subject to communication failures, it is impossible to realize an atomic read/write shared memory that guarantees a response to every request.

CAP theorem intuition

Simple situation: two interacting processes

- P and Q can no longer communicate:
  - Allow P and Q to go ahead $\Rightarrow$ no consistency
  - Allow only one of P, Q to go ahead $\Rightarrow$ no availability
- P and Q have to be assumed to continue communication $\Rightarrow$ no partitioning allowed.

CAP theorem intuition

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Fundamental question

What are the practical ramifications of the CAP theorem?

Failure detection

Issue

How can we reliably detect that a process has actually crashed?

General model

- Each process is equipped with a failure detection module
- A process P probes another process Q for a reaction
- If Q reacts: Q is considered to be alive (by P)
- If Q does not react with t time units: Q is suspected to have crashed

Observation for a synchronous system

a suspected crash $\equiv$ a known crash

Practical failure detection

Implementation

- If P did not receive heartbeat from Q within time t: P suspects Q.
- If Q later sends a message (which is received by P):
  - P stops suspecting Q
  - P increases the timeout value t
- Note: if Q did crash, P will keep suspecting Q.

Reliable remote procedure calls

What can go wrong?

1. The client is unable to locate the server.
2. The request message from the client to the server is lost.
3. The server crashes after receiving a request.
4. The reply message from the server to the client is lost.
5. The client crashes after sending a request.

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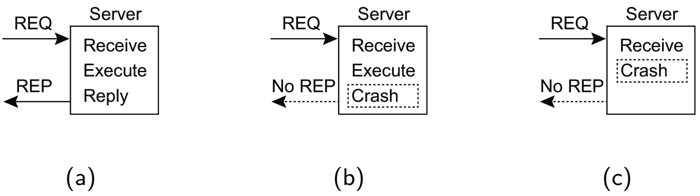
Two "easy" solutions

1: (cannot locate server): just report back to client

2: (request was lost): just resend message

Reliable RPC: server crash



Problem

Where (a) is the normal case, situations (b) and (c) require different solutions. However, we don't know what happened. Two approaches:

- At-least-once-semantics: The server guarantees it will carry out an operation at least once, no matter what.
- At-most-once-semantics: The server guarantees it will carry out an operation at most once.

Why fully transparent server recovery is impossible

Three type of events at the server

(Assume the server is requested to update a document.)

- M: send the completion message
- P: complete the processing of the document
- C: crash

Six possible orderings

(Actions between brackets never take place)

1. $M \to P \to C$: Crash after reporting completion.
2. $M \to C \to P$: Crash after reporting completion, but before the update.
3. $P \to M \to C$: Crash after reporting completion, and after the update.
4. $P \to C(\to M)$: Update took place, and then a crash.
5. $C(\to P \to M)$: Crash before doing anything
6. $C(\to M \to P)$: Crash before doing anything

Why fully transparent server recovery is impossible

Client: Reissue strategy	Server, Strategy M → P: MPC	MC(P)	C(MP)	Server, Strategy P → M: PMC	PC(M)	C(PM)
Always	DUP	OK	OK	DUP	DUP	OK
Never	OK	ZERO	ZERO	OK	OK	ZERO
Only when ACKed	DUP	OK	ZERO	DUP	OK	ZERO
Only when not ACKed	OK	ZERO	OK	OK	DUP	OK

OK = Document processed once
DUP = Document processed twice
ZERO = Document not processed at all

Reliable RPC: lost reply messages

The real issue

What the client notices, is that it is not getting an answer. However, it cannot decide whether this is caused by a lost request, a crashed server, or a lost response.

Partial solution

Design the server such that its operations are idempotent: repeating the same operation is the same as carrying it out exactly once:

- pure read operations
- strict overwrite operations

Many operations are inherently nonidempotent, such as many banking transactions.

Reliable RPC: client crash

Problem

The server is doing work and holding resources for nothing (called doing an orphan computation).

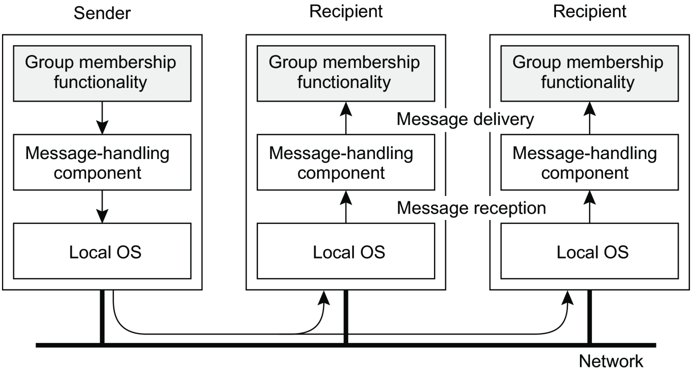
Solution

- Orphan is killed (or rolled back) by the client when it recovers
- Client broadcasts new epoch number when recovering $\Rightarrow$ server kills client's orphans
- Require computations to complete in a T time units. Old ones are simply removed.

Simple reliable group communication

Intuition

A message sent to a process group **G** should be delivered to each member of **G**.
Important: make distinction between receiving and delivering messages.



Less simple reliable group communication

Reliable communication in the presence of faulty processes

Group communication is reliable when it can be guaranteed that a message is received and subsequently delivered by all nonfaulty group members.

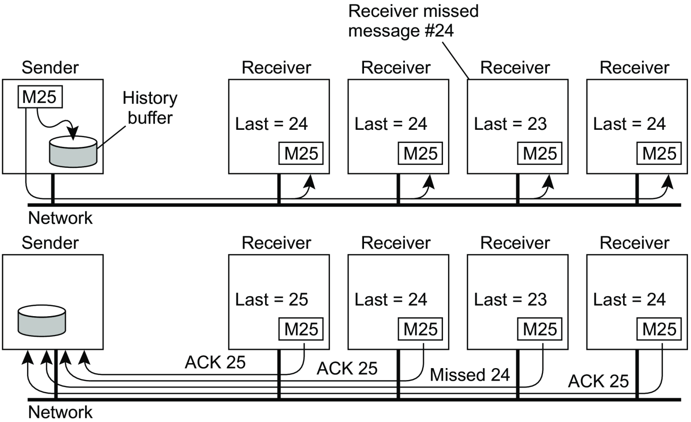
Tricky part

Agreement is needed on what the group actually looks like before a received message can be delivered.

Simple reliable group communication

Reliable communication, but assume nonfaulty processes

Reliable group communication now boils down to reliable multicasting: is a message received and delivered to each recipient, as intended by the sender.



Distributed commit protocols

Problem

Have an operation being performed by each member of a process group, or none at all.

- Reliable multicasting: a message is to be delivered to all recipients.
- Distributed transaction: each local transaction must succeed.

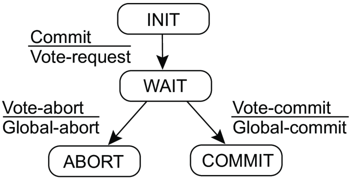
Two-phase commit protocol (2PC)

Essence

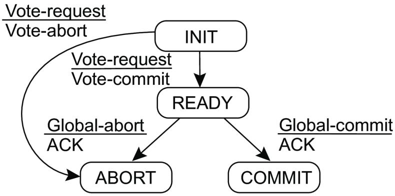
The client who initiated the computation acts as coordinator; processes required to commit are the participants.

- Phase 1a: Coordinator sends VOTE-REQUEST to participants (also called a pre-write)
- Phase 1b: When participant receives VOTE-REQUEST it returns either VOTE-COMMIT or VOTE-ABORT to coordinator. If it sends VOTE-ABORT, it aborts its local computation
- Phase 2a: Coordinator collects all votes; if all are VOTE-COMMIT, it sends GLOBAL-COMMIT to all participants, otherwise it sends GLOBAL-ABORT
- Phase 2b: Each participant waits for GLOBAL-COMMIT or GLOBAL-ABORT and handles accordingly.

2PC - Finite state machines



Coordinator



Participant

2PC – Failing participant

Analysis: participant crashes in state S, and recovers to S

- *INIT*: No problem: participant was unaware of protocol

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- *READY*: Participant is waiting to either commit or abort. After recovery, participant needs to know which state transition it should make $\Rightarrow$ log the coordinator's decision

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- *READY*: Participant is waiting to either commit or abort. After recovery, participant needs to know which state transition it should make ⇒ log the coordinator's decision
- *ABORT*: Merely make entry into abort state idempotent, e.g., removing the workspace of results
- *COMMIT*: Also make entry into commit state idempotent, e.g., copying workspace to storage.

Observation

When distributed commit is required, having participants use temporary workspaces to keep their results allows for simple recovery in the presence of failures.

2PC – Failing participant

Alternative

When a recovery is needed to *READY* state, check state of other participants ⇒ no need to log coordinator's decision.

Recovering participant P contacts another participant Q

State of Q	**Action by P**
COMMIT	Make transition to *COMMIT*
ABORT	Make transition to *ABORT*
INIT	Make transition to *ABORT*
READY	Contact another participant

Result

If all participants are in the *READY* state, the protocol blocks. Apparently, the coordinator is failing. Note: The protocol prescribes that we need the decision from the coordinator.

2PC – Failing coordinator

Observation

The real problem lies in the fact that the coordinator's final decision may not be available for some time (or actually lost).

Alternative

Let a participant P in the *READY* state timeout when it hasn't received the coordinator's decision; P tries to find out what other participants know (as discussed).

Observation

Essence of the problem is that a recovering participant cannot make a local decision: it is dependent on other (possibly failed) processes

Coordinator in Python

```python
class Coordinator:
  def run(self):
    yetToReceive = list(self.participants)
    self.log.info('WAIT')
    self.chan.sendTo(self.participants, VOTE_REQUEST)
    while len(yetToReceive) > 0:
      msg = self.chan.recvFrom(self.participants, BLOCK, TIMEOUT)
      if msg == -1 or (msg[1] == VOTE_ABORT):
        self.log.info('ABORT')
        self.chan.sendTo(self.participants, GLOBAL_ABORT)
        return
      else: # msg[1] == VOTE_COMMIT
        yetToReceive.remove(msg[0])
    self.log.info('COMMIT')
    self.chan.sendTo(self.participants, GLOBAL_COMMIT)
```

Participant in Python

```python
class Participant:
  def run(self):
    self.log.info('INIT')
    msg = self.chan.recvFrom(self.coordinator, BLOCK, TIMEOUT)
    if msg == -1:  # Crashed coordinator - give up entirely
      decision = LOCAL_ABORT
    else: # Coordinator will have sent VOTE_REQUEST
      decision = self.do_work()
      if decision == LOCAL_ABORT:
        self.chan.sendTo(self.coordinator, VOTE_ABORT)
        self.log.info('LOCAL_ABORT')
      else: # Ready to commit, enter READY state
        self.log.info('READY')
        self.chan.sendTo(self.coordinator, VOTE_COMMIT)
        msg = self.chan.recvFrom(self.coordinator, BLOCK, TIMEOUT)
        if msg == -1: # Crashed coordinator - check the others
          self.log.info('NEED_DECISION')
          self.chan.sendTo(self.participants, NEED_DECISION)
          while True:
            msg = self.chan.recvFromAny()
            if msg[1] in [GLOBAL_COMMIT, GLOBAL_ABORT, LOCAL_ABORT]:
              decision = msg[1]
              break
        else: # Coordinator came to a decision
          decision = msg[1]
    if decision == GLOBAL_COMMIT:
      self.log.info('COMMIT')
    else: # decision in [GLOBAL_ABORT, LOCAL_ABORT]:
      self.log.info('ABORT')
    while True: # Help any other participant when coordinator crashed
      msg = self.chan.recvFrom(self.participants)
      if msg[1] == NEED_DECISION:
        self.chan.sendTo([msg[0]], decision)
```

Recovery: Background

Essence

When a failure occurs, we need to bring the system into an error-free state:

- Forward error recovery: Find a new state from which the system can continue operation
- Backward error recovery: Bring the system back into a previous error-free state

Practice

Use backward error recovery, requiring that we establish recovery points

Observation

Recovery in distributed systems is complicated by the fact that processes need to cooperate in identifying a consistent state from where to recover

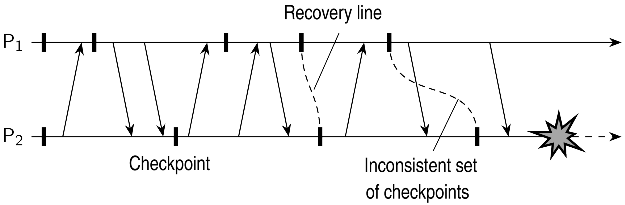
Consistent recovery state

Requirement

Every message that has been received is also shown to have been sent in the state of the sender.

Recovery line

Assuming processes regularly checkpoint their state, the most recent consistent global checkpoint.



Coordinated checkpointing

Essence

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Simple solution

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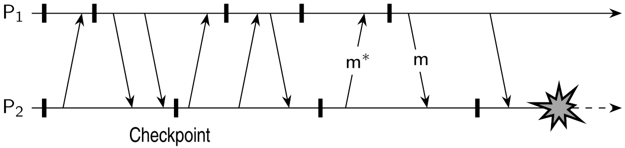
Observation

It is possible to consider only those processes that depend on the recovery of the coordinator, and ignore the rest

Cascaded rollback

Observation

If checkpointing is done at the "wrong" instants, the recovery line may lie at system startup time. We have a so-called cascaded rollback.



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Observation

If process P_i rolls back to $CP_i(m-1)$, P_j must roll back to $CP_j(n-1)$.

Message logging

Alternative

Instead of taking an (expensive) checkpoint, try to replay your (communication) behavior from the most recent checkpoint $\Rightarrow$ store messages in a log.

Assumption

We assume a piecewise deterministic execution model:

- The execution of each process can be considered as a sequence of state intervals
- Each state interval starts with a nondeterministic event (e.g., message receipt)
- Execution in a state interval is deterministic

Conclusion

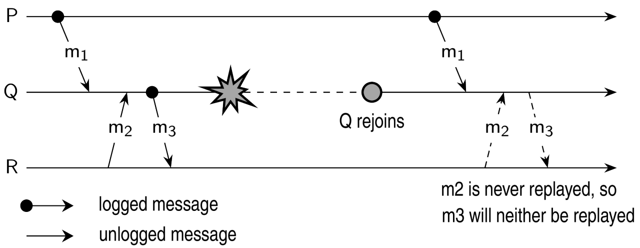
If we record nondeterministic events (to replay them later), we obtain a deterministic execution model that will allow us to do a complete replay.

Message logging and consistency

When should we actually log messages?

Avoid orphan processes:

- Process Q has just received and delivered messages m_1 and m_2
- Assume that m_2 is never logged.
- After delivering m_1 and m_2, Q sends message m_3 to process R
- Process R receives and subsequently delivers m_3: it is an orphan.



Message-logging schemes

Notations

- **DEP**(m): processes to which m has been delivered. If message m^* is causally dependent on the delivery of m, and m^* has been delivered to Q, then $Q \in \mathbf{DEP}(m)$.
- **COPY**(m): processes that have a copy of m, but have not (yet) reliably stored it.
- **FAIL**: the collection of crashed processes.

Characterization

$$Q \text{ is orphaned} \Leftrightarrow \exists m : Q \in \mathbf{DEP}(m) \text{ and } \mathbf{COPY}(m) \subseteq \mathbf{FAIL}$$

Message-logging schemes

Pessimistic protocol

For each nonstable message m, there is at most one process dependent on m, that is $|\mathbf{DEP}(m)| \leq 1$.

Consequence

An unstable message in a pessimistic protocol must be made stable before sending a next message.

Message-logging schemes

Optimistic protocol

For each unstable message m, we ensure that if $\mathbf{COPY}(m) \subseteq \mathbf{FAIL}$, then eventually also $\mathbf{DEP}(m) \subseteq \mathbf{FAIL}$.

Consequence

To guarantee that $\mathbf{DEP}(m) \subseteq \mathbf{FAIL}$, we generally roll back each orphan process Q until $Q \notin \mathbf{DEP}(m)$.