Abstract—Atomic broadcasts play a central role in serialisable in-memory transactions. Best performing ones block, when a node crashes, until a new view is installed. We augment a new protocol for uninterrupted progress in the interim period.

I. INTRODUCTION

This fast abstract is concerned with the sustainable performance of 1-copy serialisable transactions running on an in-memory database where data is partitioned and replicated over the RAM of several nodes; no node holds all replicas of a given partition nor the entire database (partial replication). In-memory databases are best suited to applications that require frequent access to large data, mainly because they offer a superior performance (e.g., by asynchronous disk writes) and can dynamically scale (by changing the number of nodes whose RAM contributions comprise the database infrastructure).

The combination of serialisability requirement and partial replication constraint poses certain challenges that are effectively addressed (see §II) by using atomic broadcasts, or abcasts for short. In brief, a transaction executes read-optimistically, get its relative order through abcast, and aborts only if its reads were made out-of-date by transactions preceding it in the abcast order; however, if the transaction is read-only and its data is replicated locally, it is exempted from the out-of-date read scrutiny. Note that while this exemption is certainly a performance enhancer, it ceases to exist whenever a transaction execution spans over items that are not stored locally as the data can be autonomously updated at other nodes.

Use of abcasts is shown [2] to be highly effective compared to the classical 2-Phase Commit approach. This comparative study uses Red Hat’s open-source in-memory database infinispan and considers all influential factors such as abort rate, latency/throughput and the average number of nodes involved in a transaction. The abcast protocol used in this study, however, is chosen from a class of protocols, such as [3], [4], that work extremely well in the absence of crashes; when a node crashes, however, they block until a group membership protocol delivers a new membership view that must also contain a virtually-synchronous closure on the set of messages that should be delivered in the old view [4]. This can take in the order of seconds, e.g., JGroups uses a default timeout of 10s to rule out false crash suspicions prior to constructing the new view. Thus, the study, albeit comprehensive otherwise, is valid only in the absence of node crashes.

The objective of our work, also sponsored by Red Hat, is to retain the best crash-free performance and to mitigate the effects of post-crash blocking. It is being accomplished in two stages: (i) delivering abcast as a separate service rather than relying on each transaction initiator itself to execute an abcast protocol with nodes involved in its transaction (§III-A); and, (ii) incorporating an insurance abcast protocol that can be effortlessly switched on or off whenever a crash is suspected or whenever a new view is ready or the suspicion has turned out to be false, respectively (§III-B). Before presenting design challenges and the achieved/expected outcomes in each stage, we highlight next the limited options available in achieving our objective.

II. APPROACH AND RATIONALE

A. No Cheap Asynchronous ABcast Insurance

Group-membership dependant (GMD for short) abcast protocols are asynchronous: they do not assume bounds on message delays nor on clock differences between nodes. When a node abcasts a message m, recipients broadcast an ack for m promising that they will only broadcast m’ with a time-stamp larger than the one m.ts found in m. When an ack is received from every member in the group, m is ready to be ordered as per m.ts. When all members are operative, a GMD abcast can have the smallest latency of one round-trip delay (when acks are broadcast instantly) and the lowest message cost of 1 broadcast (when acks are piggybacked).

The other class of asynchronous abcast protocols are quorum based (QB for short). Crash tolerance is inherent in each design/execution step: no effort is made to detect whether any node is truly crashed and measures are undertaken as though at most less than half the nodes can crash at any time soon.

Suppose that a GMD protocol is chosen as the normal abcast when no crash is suspected and a QB one as the insurance whenever a crash is suspected. Switch-over requires a virtually synchronous closure on the normal stream of already ordered, and possibly delivered, messages, i.e., constructing an agreed ‘view’ on stream closure for operative nodes is essential for consistent switch over. So, switch-over is computationally as ‘heavy-weight’ as in JGroups, except for the long 10-second duration used there to ascertain an actual crash. Using smaller timeouts can lead to false crash suspicions, making switch-over unnecessary at times. No optimal timeout exists to discern a slow node from a crashed one [5].
B. Our Approach

We use a proactive synchronous abcast protocol as the insurance. Dedicated nodes implement abcast as a service to nodes executing transactions. They keep their clocks synchronized within some known accuracy $\epsilon$ with a high probability, using \[6\]. They timestamp each $m$ and $ack$ they broadcast, which allows message delays to be pessimistically estimated. From the delays observed in the recent past, each node $i$ estimates the worst case delay $d_i$ which it encodes in its broadcast $m_i$.

When node $j$ has $m$ and is not aware of any other $m'$, $m'.ts \leq m.ts$, that is yet to be ordered, it orders $m$ (as per $m.ts$) by the GMD abcast rules or after its clock time is $m.ts + D + \epsilon$, whichever is earlier; here, $D$ is some function of $d_i$ and other parameters corresponding to various best effort protocol measures aimed at making node $j$ be aware of such $m'$. These measures are outlined in Subsection III-B.

Analysing the success of having a proactive synchronous protocol as the insurance involves two cases: node $i$ is slow or crashed and node $j$ has or knows of $m$ before $m.ts + D + \epsilon$ (Case 1) or never knows of $m$ until $m.ts + D + \epsilon$ and orders another $m'$, $m'.ts > m.ts$ (Case 2). Case 2 occurs when all best efforts within the protocol are rendered ineffective by sharp increases in communication delays. We seek to minimise the Case 2 probability to as small as $10^{-6}$.

III. CONTRIBUTIONS: COMPLETED AND EXPECTED
A. Abcast as a Service

When a node, $Tx.host$ for short, that initiated a transaction $T_x$ completes the execution, it sends a request to this (external) ordering service for a global order to be put on $T_x$, together with the list of all nodes participating in $T_x$. The request is sent to one of the multiple dedicated servers implementing the abcast service. The contacted server responds back with a global order number for $T_x$ to $Tx.host$ which, in turn, forwards the order to all participating nodes.

A subtle issue here is to ensure that a participant node $i$ is not forced to undergo a cascaded waiting when it is concurrently participating in several transactions. Say, node $i$ participates concurrently in $T_x$ and $T_x'$ initiated by $Tx.host_1$ and $Tx.host_2$, respectively. Receiving just the order number for one transaction, say, $T_x$ from $Tx.host_1$ and not (yet) for $T_x'$ does not allow node $i$ to determine if $T_x$ precedes $T_x'$ or vice versa. So, the response of an order server to a $Tx.host$ includes, for each participant node listed in the request, a short history of transactions preceding the one whose ordering has been requested. So, when node $i$ receives the order number and history for $T_x$, if $T_x'$ is not in the history for $T_x$, node $i$ can work on $T_x$, even if it has not yet received the ordering details for $T_x'$ from $Tx.host_2$.

The main advantages of abcast service are: $Tx.host$ and participant nodes are spared from executing an abcast protocol and the protocol is not restricted to be leader-based; the main cost is: time delay in contacting, and receiving the response from, the protocol. We replicated the experiments of \[2\] and the results indicate using an external order service pays off when the average number of nodes involved ($Tx.host$ and participants) exceeds 3.3. Thus, an external order service favours scalability and protocol flexibility.

B. Best Effort Design Aspects

The objectives are to (i) make a server node $i$ be aware of an order request $m$ before its clock time $m.ts + D + \epsilon$ and (ii) ensure that $D$ accommodates, as much as possible, delay variations that might occur over and above the past estimate.

On the first objective, a ‘broadcast’ of $m$ consists of two redundant broadcasts separated by some interval ($\eta$) and an $ack$ incorporates the last sequence number of the broadcast received from each server. The latter enables node $i$ to deduce any broadcast it may be missing and postpone ordering of later messages until ‘gaps’ are filled. The former enables a node to suspect that all is not well with the first broadcast, if the second broadcast has not been received within a certain timeout; it prompts a proactive response by re-broadcasting the message on behalf of the sender. (Care is taken to minimise proactive responses.) A recipient server’s response helps to complete a broadcast that may be rendered partial due to sender crash and also to fill in the ‘gaps’.

Value for $D$ is estimated as some function of $\eta$ and the probability distribution of delays estimated in the past. The function itself is designed to be pessimistic. Examples of pessimism are: a broadcast is said to be complete when the second redundant broadcast reaches recipients; the sender is always assumed to crash during the first redundant broadcast, leaving a recipient to do both the redundant broadcasts on behalf of the sender. From the cumulative distribution for $D$, we choose a value corresponding to 99.99% probability.

With the ordering service implemented by 3 dedicated server nodes, we observed no out-of-order failures for fairly-large request arrival rates. However, when arrival rates increase beyond a threshold, servers tend to saturate undermining our hypothesis that future delay can be estimated reasonably safely based on the past delay estimates. We are therefore currently implementing flow control to avoid server saturation.

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