

# Response-Time Analysis for Non-Preemptive Global Scheduling with FIFO Spin Locks

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**Abstract**—Motivated by the lack of response-time analyses for non-preemptive global scheduling that consider shared resources, this paper provides such an analysis for global job-level fixed-priority (JLFP) scheduling policies and FIFO-ordered spin locks. The proposed analysis computes response-time bounds for a set of resource-sharing jobs subject to release jitter and execution-time uncertainties by implicitly exploring all possible execution scenarios using state-abstraction and state-pruning techniques. A large-scale empirical evaluation of the proposed analysis shows it to be substantially less pessimistic than simple execution-time inflation methods, thanks to the explicit modeling of contention for shared resources and scenario-aware blocking analysis.

**Index Terms**—real-time systems, response-time analysis, shared resources, global multiprocessor scheduling

## I. INTRODUCTION

An essential capability provided by virtually all multitasking real-time operating systems (or runtime environments) is the ability to share data or software resources without running into race conditions. Multiprocessor real-time systems typically use two types of locking protocols for this purpose: suspension-based and spin-based protocols, which offer various tradeoffs. Although the literature provides ample analyses of both of these approaches for globally scheduled *preemptive* tasks (e.g., see a recent survey [10]), to date only little attention has been paid to high-accuracy response-time analysis for globally scheduled *non-preemptive* tasks that share resources.

To close this gap, we propose a blocking-aware response-time analysis for non-preemptive tasks that are scheduled by a work-conserving global *job-level fixed-priority* (JLFP) scheduling policy on a multicore platform. We focus on spin-lock protocols since they are a natural fit for non-preemptive scheduling: they are relatively simple to implement, require virtually no OS support, and are compatible with run-to-completion (or single-stack) runtime systems. Furthermore, spin locks incur significantly lower runtime overhead (relative to suspension-based locking protocols) [9], so that the cost of suspending and resuming a task can easily outweigh the cost of busy-waiting (i.e., spinning) if critical sections are relatively short (which they usually are in well-designed systems). Finally, while it is notoriously difficult to predict cache contents after suspensions (usually, it is pessimistically assumed that any suspension results in a complete loss of cache affinity), a spinning job effectively *protects* the state of its assigned processor and cache. This makes the system not only more

time-predictable, but also goes a long way towards enabling a sound *worst-case execution time* (WCET) analysis.

More specifically, in **this paper**, we propose a new schedulability analysis for periodic tasks and other workloads with a repeating pattern of job releases (such as multi-frame tasks). This choice is motivated by the fact that many practical workloads are periodic (e.g., in automotive systems, in control applications, in video and audio processing, in computer vision etc.). In fact, a recent survey of 120 industry practitioners by Akesson et al. [4] indicated that 82% of the respondents work on systems that include periodic tasks, and 21% of the respondents answered that their system is exclusively comprised of tasks with periodic and/or table-driven activation.

The analysis proposed in this paper is based on the notion of schedule abstraction, a concept initially introduced by Nasri et al. [21]–[23]. However, whereas Nasri et al.’s prior analyses do not consider locking-induced delays, we assume—and model in detail—that tasks coordinate mutually exclusive access to shared resources by means of *FIFO-ordered spin locks*, a common choice in multiprocessor real-time systems [10]. Our analysis is generic in nature and covers all work-conserving global non-preemptive JLFP schedulers such as *non-preemptive earliest-deadline first* (NP-EDF) and *non-preemptive fixed-priority* (NP-FP) scheduling. Furthermore, our analysis is *sufficient*, i.e., it covers any sequence of job finish times that may occur in the modeled system.

The proposed analysis implicitly explores all possible orders of job start times as well as their accesses to shared resources in a *schedule-abstraction graph* [21]–[23]. Each such ordering results in a different sequence of *system states*. The efficiency of the analysis (in terms of runtime and memory footprint) and accuracy of this exploration strongly depends on the level of abstraction used to encode the system states. To this end, we propose a completely *new system-state abstraction* that accurately models shared resource accesses by means of FIFO spin locks. This new system state representation requires the design (and a proof of soundness of) a whole new set of expansion and merging rules, which are used to build the schedule-abstraction graph. Ultimately, we prove in Section IV-I that all possible system states are indeed explored and that our analysis yields a safe *worst-case response time* (WCRT) bound for each job, which reflects the worst-case blocking suffered by those jobs.

Notably, our analysis does *not* just pessimistically account

for the *union* of all critical sections that could theoretically overlap with a job in *some* schedule; rather, our analysis is *execution-scenario-aware* and accounts only for contention that can actually arise in a specific interleaving of resource accesses by the jobs. For example, if a job incurs *either* shared-resource contention (if it is scheduled right away) *or* interference from higher-priority jobs (but not both), then the proposed analysis will reflect the worse of the two scenarios, but *not* an infeasible combination of both (scheduling interference *and* resource contention). We are not aware of any prior response-time analysis with a comparable level of accuracy in the accounting of synchronization delays.

In fact, somewhat surprisingly, we found that there is hardly any directly related **prior work** [10]. While FIFO-ordered spin locks in real-time systems have long been recognized to have favorable properties for worst-case analysis [3,5]–[7,9,12,13, 15,18,20,27], prior work in this space has focused on either *preemptive* global scheduling [3,7,9,12,15,20] or partitioned scheduling [5]–[7,9,18,27].

In the context of preemptive global scheduling, the most relevant locking protocols and blocking analyses are Brandenburg’s *holistic blocking analysis* of non-preemptive FIFO spin locks [9] as used in Block et al.’s FMLP for *short* critical sections [7], and the *suspension-based* global OMLP [11], FMLP for *long* critical sections [7,28], and PIP [16,25,28]. We compare against each of these in our evaluation (Sec. V).

For a review of real-time locking in general, we refer the reader to Brandenburg’s recent comprehensive survey of the area [10].

Finally, we note that, while it might be tempting to compare the timing analysis of FIFO-ordered spin lock accesses to queuing theory or the analysis of networks, our given problem is fundamentally different from packets traversing a network since dealing with locks in real-time systems is essentially a multi-resource scheduling problem (i.e., jobs need to get access to both the processor and the shared resources at the same time but for different durations). More precisely, we cannot solely model our problem as a network of queues to be traversed because the queuing delays for access to processors and shared resources interact in nontrivial ways. This makes a direct comparison to the domain of queuing theory, Real-Time Calculus [26], or Network Calculus [8] inapplicable.

To our knowledge, there is no prior work on the central contribution of this paper: a *blocking-aware* response-time analysis for *non-preemptive* global scheduling with shared resources protected by FIFO-ordered spin locks.

## II. SYSTEM MODEL AND DEFINITIONS

We consider a work-conserving global JLFP scheduling policy to schedule a set of *non-preemptive* tasks on a multicore platform with  $m$  identical cores. Each task releases jobs according to a specific pattern (e.g., periodic, rate-based or bursty), and each job has a fixed priority. The JLFP scheduler dispatches ready jobs in order of decreasing priority.

### A. Workload Model

A job  $J_i = ([r_i^{min}, r_i^{max}], d_i, p_i, \langle J_{i,1}, \dots, J_{i,n_i} \rangle)$  has an earliest release time  $r_i^{min}$ , a latest release time  $r_i^{max}$ , an absolute deadline  $d_i$ , and a priority  $p_i$ . We assume that timing parameters are discrete, i.e., integer multiples of a clock tick.

A job’s execution is modeled by its sequence of *job segments*  $\langle J_{i,1}, \dots, J_{i,n_i} \rangle$ , where  $J_{i,1}$  and  $J_{i,n_i}$  are its first and last segment, respectively. Each job segment  $J_{i,j}$  is identified by  $J_{i,j} = ([C_{i,j}^{min}, C_{i,j}^{max}], \eta_{i,j}, [L_{i,j}^{min}, L_{i,j}^{max}])$ , where  $C_{i,j}^{min}$  is the best-case execution time (BCET),  $C_{i,j}^{max}$  is the worst-case execution time (WCET),  $\eta_{i,j}$  is the set of resources that the segment *requests*, and  $L_{i,j}^{min}$  and  $L_{i,j}^{max}$  are the minimum and maximum length of the critical section associated with  $\eta_{i,j}$ , respectively. We obviously must have that  $L_{i,j}^{min} \leq C_{i,j}^{min}$  and  $L_{i,j}^{max} \leq C_{i,j}^{max}$  since the critical section length is part of the execution time of the segment. In this work, we assume that each segment  $J_{i,j}$  can access at most one resource protected by a lock (i.e.,  $|\eta_{i,j}|$  is equal to 0 or 1). Without any loss of generality, we further assume that the critical section of a segment is located at the beginning of the segment. The segment must first be granted the lock protecting the shared resource in  $\eta_{i,j}$  (if any) to start executing.

All segments of a job inherit the job’s priority. For ease of notation, we use  $hp(J_{i,j})$  to refer to the set of segments with a higher priority than  $J_{i,j}$ . We assume that ties in priority are broken arbitrarily but consistently.

Each job must execute sequentially (on one core) with run-to-completion semantics, but may compete for shared resources with other jobs running in parallel on other cores. We do not consider inter-job precedence constraints in this paper.

### B. Shared Resources

We let  $\mathcal{L}$  denote the set of shared resources protected by locks. A shared resource  $\ell_x \in \mathcal{L}$  is *available* iff no job is currently accessing it. When a job segment  $J_{i,j}$  is given access to a resource, we say that the resource is *granted* to  $J_{i,j}$ . A segment  $J_{i,j}$  cannot start executing until the resource  $\ell_x \in \eta_{i,j}$  (if any) is granted to  $J_{i,j}$ . Resource  $\ell_x$  will be released as soon as the segment’s critical section completes (i.e., within  $[L_{i,j}^{min}, L_{i,j}^{max}]$  time units after the resource was granted). If a segment  $J_{i,j}$  requests access to a resource  $\ell_x$  that is already granted to another job, we say that  $J_i$  is *blocked* on  $\ell_x$ . The job  $J_i$  *busy-waits* (i.e., *spins*) until it is granted  $\ell_x$ .

In this paper, we assume FIFO spin locks, i.e., resources are granted in the order in which requests arrive. If two jobs request the same resource at the exact same time, we assume that the tie is broken arbitrarily but consistently (e.g., in our experiments, ties are broken in favor of lower job IDs).

### C. Execution Model

A job is *ready* at time  $t$  if it has been released and has not yet started executing. Due to the non-preemptive execution model, a job must run to completion without preemption once it has started execution. Thus, once the first segment  $J_{i,1}$  of a job  $J_i$  begins its execution on a core, the core is exclusively serving  $J_i$  until  $J_i$  completes its execution, including any blocking

delays. We say that a job  $J_i$  has *claimed* a core if it started executing and has not finished yet. Upon completion we say that  $J_i$  *releases* the core, making it again available to other jobs. Hence, a core is either *claimed* (busy executing a job) or *free* for a new job to start executing.

Given the *finish time*  $f_i$  and the earliest-release time  $r_i^{min}$  of  $J_i$ , we compute the *response time* of the job as  $f_i - r_i^{min}$ .

#### D. Problem Definition

We seek to bound the worst-case response time of each member of a finite set of non-preemptive jobs  $\mathcal{J}$  that access shared resources. The job set  $\mathcal{J}$  is the set of all jobs released in an *a priori* computed *observation window*. The input job set in the observation window must be a representative workload for the system, namely, either because the observation window is long enough to capture the whole workload (e.g., in batch scheduling), or because the release patterns of the jobs in the observation window repeats for the rest of the lifetime of the system so that it is sufficient to only analyze one instance. For example, the observation window of synchronous periodic task sets with implicit deadlines that exhibit no deadline misses is one *hyperperiod* (i.e., the least integer multiple of the periods); further safe observation windows for various types of periodic (or multi-frame) task models can be found in [19,21,24].

Our analysis deems a job set  $\mathcal{J}$  *schedulable* only if no execution scenario of  $\mathcal{J}$  (Definition 1 below) leads to some job exhibiting a response time exceeding its deadline. We extend the original definition of execution scenario [21] to reflect shared-resource accesses.

**Definition 1.** An execution scenario  $\gamma$  is a mapping of jobs to release times, segment execution times, and critical-section lengths such that  $\forall J_i \in \mathcal{J}$ ,  $r_i \in [r_i^{min}, r_i^{max}]$  and  $\forall j, 1 \leq j \leq n_i$ ,  $C_{i,j} \in [C_{i,j}^{min}, C_{i,j}^{max}]$  and  $L_{i,j} \in [L_{i,j}^{min}, L_{i,j}^{max}]$ .

### III. SCHEDULABILITY ANALYSIS

The proposed schedulability analysis builds a *schedule-abstraction graph* (SAG) [21] to implicitly explore all possible execution scenarios of a job set. In this section, we define our notion of a *resource-aware* SAG and explain its construction and use in bounding each job's worst-case response time.

#### A. Schedule-Abstraction Graph

A SAG is a directed acyclic graph  $G = (V, E)$ , where  $V$  indicates the set of (abstract) states reachable by the system and  $E$  represents the set of scheduling decisions leading from one (abstract) system state to another. An edge  $e = (v_p, v_q, J_{i,j})$  from  $v_p$  to  $v_q$  with label  $J_{i,j}$  indicates that executing  $J_{i,j}$  in state  $v_p$  evolves the system to state  $v_q$ . We say that a job segment  $J_{i,j}$  starts executing *next* in state  $v_p$  (or *succeeds*  $v_p$ ) if there exists an outgoing edge from  $v_p$  with label  $J_{i,j}$ .

By convention [21], state  $v_1$  represents the initial state of the system at time zero, where every core is idle and no segment has started executing yet. A path  $P$  from  $v_1$  to a state  $v_p$  represents a possible job-segment start order that results in system state  $v_p$ . The length of such a path  $P$  indicates the number of segments that have started (and potentially already

finished) their execution when the system reaches state  $v_p$ , i.e.,  $|P| \triangleq |\mathcal{J}^P|$ , where  $\mathcal{J}^P$  denotes the set of labels on the edges of path  $P$ . If a vertex  $v_p$  has multiple incoming edges, then the scheduling decisions that lead to  $v_p$  must involve the same set of job segments on any two paths from  $v_1$  to  $v_p$ .

**Property 1.** (adapted from [23]) For any two paths  $P$  and  $Q$ , if both  $P$  and  $Q$  start at  $v_1$  and end in  $v_p$ , then  $\mathcal{J}^P = \mathcal{J}^Q$ .

#### B. System State Representation

As previously stated, we consider that jobs can be subject to release jitter and that there can be variations in the execution times of its segments. As a consequence of this uncertainty, we must compute a finish time interval  $[EFT_{i,j}, LFT_{i,j}]$  for each segment of a job  $J_i$  after a sequence of scheduling decisions taken by the scheduler. This interval is bounded by the job segment's earliest and latest finish times  $EFT_{i,j}$  and  $LFT_{i,j}$  in any execution scenario that complies with the assumed sequence of scheduling decisions. Thus, we say that a job segment  $J_{i,j}$  of job  $J_i$  can possibly finish at or after  $EFT_{i,j}$  and is certainly finished by  $LFT_{i,j}$ . This uncertainty in the finish times of segments introduces a challenge as it also means that we have uncertainty in processor and shared-resource availability times, which then affect the finish time intervals of subsequent job segments.

To address this challenge, we develop a new abstraction to encode information about the system state. In each vertex  $v_p$ , the system state abstraction records the availability of the cores and shared resources for any execution scenario that complies with the sequence of scheduling decisions modeled by the edges on the paths leading to  $v_p$ .

As discussed in Section II-C, a core is either free to execute a ready job or claimed by a job that was previously dispatched by the scheduler, i.e., a job has started executing on the core and not all its segments have finished executing yet. Thus, let  $\mathcal{C}(v_p)$  denote the set of jobs that claimed a core (where  $|\mathcal{C}(v_p)| \leq m$ ). A system state is then modeled as follows:

- **Claimed core availability intervals.** For each job  $J_i \in \mathcal{C}(v_p)$ , we record the interval  $[Cl_i^{min}(v_p), Cl_i^{max}(v_p)]$ , where  $Cl_i^{min}(v_p)$  and  $Cl_i^{max}(v_p)$  indicate when the core claimed by job  $J_i$  becomes *possibly* and *certainly* available to execute the next segment of  $J_i$ , respectively.
- **Free-core availability intervals.** We record  $m - |\mathcal{C}(v_p)|$  intervals  $A_x(v_p) = [A_x^{min}(v_p), A_x^{max}(v_p)]$  such that  $1 \leq x \leq m - |\mathcal{C}(v_p)|$ , which indicate when one, two, three, ...,  $m - |\mathcal{C}(v_p)|$  free cores (i.e., those that are not already claimed by executing jobs) become *possibly* and *certainly* available to execute *ready jobs*.
- **Shared-resource availability intervals.** For each shared resource  $\ell_x \in \mathcal{L}$ , a state records the interval  $[SR_x^{min}(v_p), SR_x^{max}(v_p)]$ , where  $SR_x^{min}(v_p)$  and  $SR_x^{max}(v_p)$  denote the times at which  $\ell_x$  become *possibly* and *certainly* available again, respectively.

For ease of reference, we also introduce the two notations  $SR_{i,j}^{min}(v_p)$  and  $SR_{i,j}^{max}(v_p)$  to refer to the availability interval of the resource accessed by segment  $J_{i,j}$  (if any).

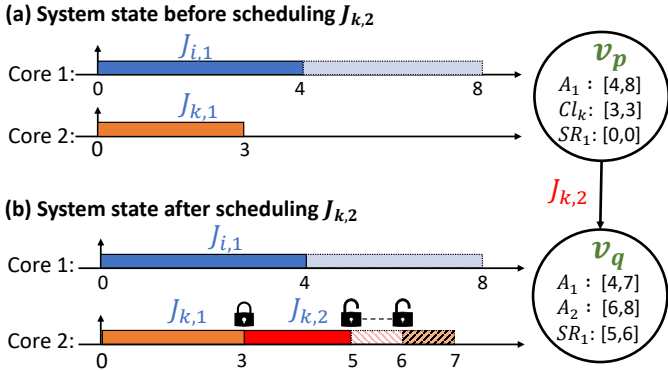


Fig. 1: Two successive system states for  $m = 2$  and two jobs  $J_i$  (with one segment) and  $J_k$  (with two segments) released at time 0. Only the second segment of  $J_k$  requests a shared resource  $\ell_1$  with a critical section length  $L_{k,2} \in [2, 3]$ . The execution times of the segments are  $C_{i,1} \in [4, 8]$ ,  $C_{k,1} \in [3, 3]$  and  $C_{k,2} \in [3, 4]$ .

Hence, if  $J_{i,j}$  accesses a shared resource (i.e.,  $\exists \ell_x \in \eta_{i,j}$ ), then  $SR_{i,j}^{min}(v_p)$  and  $SR_{i,j}^{max}(v_p)$  are the earliest and latest availability time  $SR_x^{min}(v_p)$  and  $SR_x^{max}(v_p)$  of the resource  $\ell_x$  accessed by  $J_{i,j}$ . If  $J_{i,j}$  does not access a resource, then simply  $SR_{i,j}^{min}(v_p) = SR_{i,j}^{max}(v_p) = 0$ .

**Example.** Fig. 1 shows how an abstract state  $v_p$  in the SAG evolves into a new state  $v_q$  when a ready job segment  $J_{k,2}$  is scheduled. On the left, we show the status of the cores before and after scheduling  $J_{k,2}$ . The right side shows the contents of the abstract state  $v_p$  and how it evolves after scheduling  $J_{k,2}$ .

Since  $J_i$  has only one segment, its start time defines its finish time interval and thus the time at which Core 1 may become available for ready jobs (between times 4 and 8). Thus,  $v_p$  records that one core is free for a ready job to execute within the interval  $[4, 8]$  (i.e.,  $A_1(v_p) = [4, 8]$ ). State  $v_p$  also records that one core is claimed by job  $J_k$  whose first segment  $J_{k,1}$  will potentially and certainly complete by time 3 (i.e.,  $Cl_k = [3, 3]$ ). Moreover, since none of the two segments executing in  $v_p$  access a shared resource, the shared resource  $\ell_1$  is certainly available from time 0 (i.e.,  $SR_1 = [0, 0]$ ).

Prior to executing  $J_{k,2}$ , the lock protecting the shared resource it requires must be acquired. Since the shared resource is available, the resource is granted to  $J_{k,2}$  as soon as the core already claimed by job  $J_k$  becomes available, i.e., at time 3. Since  $J_{k,2}$ 's critical section has a variable execution time, it may release the lock any time in the interval  $[5, 6]$ . This is recorded in the state  $v_q$  with the interval  $SR_1 = [5, 6]$ . Moreover, since  $J_{k,2}$  is the last segment of its job, as soon as it finishes, Core 2 is released and becomes available for other jobs. Hence, in state  $v_q$ , no core is claimed anymore. Since the execution time of  $J_{k,2}$  ranges from 3 to 4 time units, the core claimed by  $J_{k,2}$  will become available in the interval  $[6, 7]$ . This means that one core becomes possibly available at time 4 (Core 1), one core is certainly available at time 7 (Core 2), two cores are possibly available at time 6, and two cores are certainly available at time 8. Thus,  $v_q$  records the

### Algorithm 1: Construction of the SAG for job set $\mathcal{J}$ .

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**Inputs :** Job set  $\mathcal{J}$   
**Output:** Bounds on the BCRT and WCRT of every job in  $\mathcal{J}$

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1  $\forall J \in \mathcal{J}, BR_i \leftarrow \infty, WR_i \leftarrow 0$  ;
2 Initialize  $G$  by adding  $v_1 = (\{[0, 0], \dots, [0, 0]\}, \{[0, 0], \dots, [0, 0]\})$  ;
3 while  $\exists$  path  $P$  from  $v_1$  to a leaf vertex such that  $P \neq \mathcal{J}$  do
4    $P \leftarrow$  the shortest path from  $v_1$  to a leaf vertex  $v_p$ ;
5    $\mathcal{R}^P \leftarrow$  set of potentially ready segments obtained with Eq. (1);
6   for each segment  $J_{i,j} \in \mathcal{R}^P$  do
7     if  $J_{i,j}$  can start executing after  $v_p$  (Theorem 1) then
8       Build  $v'_p$  using Algorithm 2;
9       if  $J_{i,j} = J_{i,n_i}$  then
10         $BR_i \leftarrow \min\{EFT(v'_p) - r_i^{min}, BR_i\}$  ;
11         $WR_i \leftarrow \max\{LFT(v'_p) - r_i^{max}, WR_i\}$  ;
12      end
13      Connect  $v_p$  to  $v'_p$  with an edge labeled  $J_{i,j}$ ;
14      if  $\exists$  path  $Q$  that ends in  $v_q$  such that Rule 1 is
15        satisfied for  $v'_p$  and  $v_q$  then
16        Merge  $v'_p$  and  $v_q$  in  $v_z$  using Eq. (16)-(18);
17        Redirect all incoming edges of  $v_q$  and  $v'_p$  to  $v_z$ ;
18        Remove  $v_q$  and  $v'_p$  from  $G$ ;
19      end
20    end
21 end

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availability intervals  $A_1(v_q) = [4, 7]$  and  $A_2(v_q) = [6, 8]$ .

Note that every abstract system state records exactly  $m$  core availability intervals. They are split in  $|\mathcal{C}(v_p)|$  intervals for claimed cores (those on which the first segment of a job was dispatched and its last segment was not dispatched yet) and  $m - |\mathcal{C}(v_p)|$  intervals for free cores (i.e., any core that is not claimed). Therefore, the number of claimed and free cores does not depend on a specific time point but rather on the specific set of in-progress segments.

### C. Schedule-Abstraction Graph Generation

The SAG is built iteratively. Each iteration comprises two phases, namely the *expansion phase* and the *merge phase*. In the expansion phase, (one of) the shortest path(s)  $P$  in the graph is picked and expanded for every job segment that can possibly start executing next in the state  $v_p$  at the end of  $P$ . For every such job segment  $J_{i,j}$ , we create a new vertex  $v'_p$  in the graph, which represents a new system state and is connected to  $v_p$  via a directed edge labeled with  $J_{i,j}$ . The information encoded in the new state  $v'_p$  contains an updated version of the free cores, claimed cores and shared-resource availability intervals reflecting that  $J_{i,j}$  has now started to execute.

After the graph has been expanded with new states, the merge phase commences. The purpose of the merge phase is to slow down the growth of the graph, to postpone a potential state-space explosion for as long as possible. This is achieved by merging any two “similar” states that terminate paths with identical sets of job segments. To preserve soundness, every system state that is reachable from any of the two original states must also be reachable from the merged state, which ensures that no possible execution scenario is discarded.

The exploration completes when no vertex remains to be expanded, i.e., all job segments have been scheduled on every

path. At this point, each path represents a set of valid execution scenarios and every possible schedule has been explored.

The complete SAG construction procedure is given in Algorithm 1. Note that it uses two variables ( $BR_i$  and  $WR_i$ ) for each job in  $\mathcal{J}$  to keep track of its best-case (BCRT) and worst-case response time (WCRT) on any path in the graph. Those bounds are updated (lines 10–11) whenever the last segment of a job is scheduled on a path, hence indicating the completion of the job in that execution scenario. If, by the end of the exploration, no job has a WCRT exceeding its deadline, then the analysis deems the job set schedulable.

#### IV. EXPANSION PHASE

We now consider the expansion and merge phases and show how a path ending in state  $v_p$  is expanded to a new state  $v_q$ .

##### A. Overview

First, we build a set of potentially *ready* job segments in state  $v_p$ , i.e., the segments that have not started executing on path  $P$  and for which all preceding segments have started and potentially completed. For each such segment  $J_{i,j}$ , we compute the earliest and latest time  $EST_{i,j}(v_p)$  and  $LST_{i,j}(v_p)$  at which the segment can start executing in  $v_p$ . That is,  $EST(v_p)$  denotes the earliest time at which a global work-conserving JLFP scheduler would allow  $J_{i,j}$  to start in  $v_p$ . Similarly,  $LST_{i,j}(v_p)$  indicates the latest time at which  $J_{i,j}$  must have certainly started executing if it is the next segment to succeed  $v_p$ . If  $J_{i,j}$  has not started by  $LST_{i,j}(v_p)$ , a different segment must have started to execute after  $v_p$ . Hence, a segment is said to be eligible to be a successor of  $v_p$  only if its earliest start time  $EST_{i,j}(v_p)$  is no later than its latest start time  $LST_{i,j}(v_p)$ . For each eligible job segment, a new vertex  $v'_p$  is added to the graph, where  $v'_p$  encodes the system state after  $J_{i,j}$  started executing. In addition to deciding whether a job segment is eligible to start executing after state  $v_p$ ,  $EST_{i,j}(v_p)$  and  $LST_{i,j}(v_p)$  also help compute the earliest and latest finish times  $EFT_{i,j}(v_p)$  and  $LFT_{i,j}(v_p)$  of  $J_{i,j}$ .

Next, each step of the expansion phase is explained in detail.

##### B. Ready Interval

We consider that a job segment is ready if the job has been released and all its preceding segments have completed. Thus, we define the set of *potentially ready* segments  $\mathcal{R}^P$  for path  $P$  to be the set of segments that have not yet started to execute (i.e.,  $J_{i,j} \notin \mathcal{J}^P$ ) and whose predecessor (if any) has already started and potentially completed its execution (i.e., either  $J_{i,j}$  is the first segment of its job  $J_i$ , thereby meaning  $j = 1$ , or  $J_{i,j-1} \subseteq \mathcal{J}^P$ ).

$$\mathcal{R}^P \triangleq \{J_{i,j} \mid J_{i,j} \notin \mathcal{J}^P \wedge (j = 1 \vee J_{i,j-1} \subseteq \mathcal{J}^P)\} \quad (1)$$

##### C. Earliest Start Time

Since we consider that jobs may suffer from release jitter and execution-time variation, the exact finishing times of preceding job segments is unknown. Therefore, the exact time at which a segment  $J_{i,j}$  may start to execute is also unknown. For each segment  $J_{i,j}$  in

$\mathcal{R}^P$ , we prove a lower bound  $EST_{i,j}(v_p)$  and an upper bound  $LST_{i,j}(v_p)$  on the time at which  $J_{i,j}$  may start executing in  $v_p$  (Equations (2) and (7), respectively).

$$EST_{i,j}(v_p) = \begin{cases} \infty & \text{if } j = 1 \wedge |\mathcal{C}(v_p)| = m \\ \max\{r_i^{min}, A_1^{min}(v_p), SR_{i,j}^{min}(v_p)\} & \text{if } j = 1 \wedge |\mathcal{C}(v_p)| < m \\ \max\{Cl_i^{min}(v_p), SR_{i,j}^{min}(v_p)\} & \text{if } j > 1 \end{cases} \quad (2)$$

**Lemma 1.** *Segment  $J_{i,j} \in \mathcal{R}^P$  cannot start executing (as a successor of state  $v_p$ ) before  $EST_{i,j}(v_p)$ .*

*Proof:* The first segment  $J_{i,1}$  of a job  $J_i$  can start its execution only if (i) it is released, (ii) the shared resource  $\ell_x$  it requests (if any) is available **and** (iii) a core is available. Thus, if all cores have already been claimed by other jobs (i.e.,  $|\mathcal{C}(v_p)| = m$ ) then  $J_{i,1}$  cannot be a successor of  $v_p$  and  $EST_{i,j}(v_p) = \infty$ . This proves the first case of Eq. (2).

However, if there is a free core (i.e.,  $|\mathcal{C}(v_p)| < m$ ), then, by definition,  $A_1^{min}$  is the earliest time at which a core can potentially become available. Furthermore,  $r_i^{min}$  is the earliest release time of  $J_i$  and  $SR_{i,j}^{min}(v_p)$  is the earliest time at which the shared resource accessed by  $J_{i,j}$  may become available. Thus, the earliest time at which  $J_{i,1}$  may start to execute is given by  $EST_{i,j}(v_p) = \max\{r_i^{min}, A_1^{min}(v_p), SR_{i,j}^{min}(v_p)\}$  if  $j = 1$ . This proves the second case of Eq. (2).

Any segment that is not the first segment of a job (i.e., a segment  $J_{i,j}$  with  $j > 1$ ) can start its execution only if (i) the core claimed by the preceding segments belonging to the same job is available, **and** (ii) the shared resource it requests (if any) is also available. Since  $J_{i,j}$  is in  $\mathcal{R}^P$ , all the segments of  $J_i$  that precede  $J_{i,j}$  must have started (and potentially finished) executing on the core claimed by  $J_i$ . Thus, the earliest time at which the core claimed by  $J_i$  may become available for  $J_{i,j}$  is given by  $Cl_i^{min}(v_p)$ . Therefore, the earliest time at which  $J_{i,j}$  may start to execute is  $\max\{Cl_i^{min}(v_p), SR_{i,j}^{min}(v_p)\}$  if  $j > 1$ . This proves the last case of Eq. (2). ■

##### D. Latest Start Time

The latest start time  $LST_{i,j}(v_p)$  of segment  $J_{i,j}$  is computed considering that the scheduler is (i) work-conserving and (ii) that it follows a non-preemptive JLFP scheduling policy.

First, focus on (i). By definition, a work-conserving scheduler *must* execute a segment as soon as the segment is certainly ready and a core is certainly available. We can thus prove the following two lemmas and their corollary.

**Lemma 2.** *An upper bound on the time at which a segment  $J_{y,z}$  can certainly start executing (as a successor of state  $v_p$ ) is given by*

$$tw_{y,z} = \begin{cases} \infty & \text{if } z = 1 \wedge |\mathcal{C}(v_p)| = m; \\ \max\{r_y^{max}, A_1^{max}, SR_{y,z}^{max}(v_p)\} & \text{if } z = 1 \wedge |\mathcal{C}(v_p)| < m; \\ \max\{Cl_y^{max}(v_p), SR_{y,z}^{max}(v_p)\} & \text{if } z > 1. \end{cases} \quad (3)$$

*Proof:* Infinity is an obvious upper bound on the start time of  $J_{i,j}$ . Therefore, in this proof, we only focus on the cases where  $tw_{y,z} \neq \infty$ .

A starting segment  $J_{y,1}$  must start to execute as soon as (i) it is released, (ii) the resource  $\ell_x$  it requests (if any) is available and (iii) a core is available. By definition,  $r_y^{max}$  is an upper bound on the release time of  $J_{y,1}$ ,  $SR_{y,z}^{max}(v_p)$  is an upper bound on the availability time of the shared resource accessed by  $J_{y,z}$  (if any) and  $A_1^{max}(v_p)$  denotes the time at which a core is certainly available in state  $v_p$ . Thus,  $J_{y,1}$  can certainly start executing at  $\max\{r_y^{max}, A_1^{max}, SR_{y,1}^{max}(v_p)\}$  when  $z = 1$  and there is at least one free core in  $v_p$ .

Any segment  $J_{y,z}$  that is not the first segment of  $J_y$  (i.e., with  $z > 1$ ) can certainly start executing when (i) the resource it requests (if any) is available, and (ii) the segments of  $J_y$  that precede  $J_{y,z}$  have all completed their execution on the core claimed by  $J_y$ . Since  $J_{y,z}$  is in  $\mathcal{R}^P$ , all the segments of  $J_{y,z}$  preceding  $J_{y,z}$  have already started (and potentially finished) executing on the core claimed by  $J_y$ , and because  $Cl_y^{max}(v_p)$  is an upper bound on the time at which the core claimed by  $J_y$  becomes available to execute the next segment of  $J_y$ , we have that  $J_{y,z}$  certainly starts at  $\max\{Cl_y^{max}(v_p), SR_{y,z}^{max}(v_p)\}$ . ■

**Lemma 3.** *A work-conserving scheduler will start executing a job segment no later than<sup>1</sup>*

$$t_{wc} = \min_{\infty}\{tw_{y,z} \mid J_{y,z} \in \mathcal{R}^P\} \quad (4)$$

*Proof:* By Lemma 2, a job segment  $J_{y,z} \in \mathcal{R}^P$  is certainly ready to start executing at time  $tw_{y,z}$ . Thus there is at least one job segment that is ready to execute at time  $\min_{\infty}\{tw_{y,z} \mid J_{y,z} \in \mathcal{R}^P\}$ . Thus,  $t_{wc}$  is an upper-bound on the time at which a work-conserving scheduler will start executing a job segment. ■

**Corollary 1.** *A segment  $J_{i,j} \in \mathcal{R}^P$  cannot be a direct successor of a state  $v_p$  if it starts executing any later than  $t_{wc}$ .*

*Proof:* Since the scheduler will certainly schedule a job segment by time  $t_{wc}$ ,  $J_{i,j}$  must start executing at or before  $t_{wc}$  if it is the first job scheduled after  $v_p$ . ■

Now that we covered work conservation, we consider the impact of the JLFP scheduling policy.

First, assume that two segments  $J_{i,j}$  and  $J_{y,z}$  request the same resource. Since a FIFO spin lock provides access to the shared resource based on the order in which requests were made, the order in which those segments will start executing does not depend on their priority but on the order of their releases instead. Therefore, the JLFP scheduling policy does not impact the start time of segments that share the same resource. Then, let  $\mathcal{H}$  denote the set of segments with a higher priority than  $J_{i,j}$  that do not share a resource with  $J_{i,j}$ , i.e.,  $\mathcal{H} = \{J_{y,z} \mid J_{y,z} \in \{\mathcal{R}^P \cap hp(J_{i,j})\} \wedge \eta_{y,z} \cap \eta_{i,j} = \emptyset\}$ . Let  $t_{high}$  be an upper-bound on the time at which any segment in

$\mathcal{H}$  will certainly be ready to execute. We prove (Lemmas 4 and 5) that  $t_{high}$  can be computed with Eq. (5).

$$t_{high} = \min_{\infty}\{th_{y,z}(J_{i,j}) \mid J_{y,z} \in \mathcal{H}\} \quad (5)$$

where

$$th_{y,z}(J_{i,j}) = \begin{cases} \max\{r_y^{max}, SR_{y,z}^{max}(v_p)\} & \text{if } z = j = 1 \\ \max\{A_1^{max}(v_p), r_y^{max}, SR_{y,z}^{max}(v_p)\} & \text{if } z = 1 \wedge j > 1 \\ \max\{Cl_y^{max}(v_p), SR_{y,z}^{max}(v_p)\} & \text{if } z > 1 \end{cases} \quad (6)$$

**Lemma 4.** *Let  $J_{y,z}$  be a segment in  $\mathcal{H}$ . If  $J_{i,j}$  did not start to execute before  $th_{y,z}(J_{i,j})$ , then  $J_{y,z}$  will start before  $J_{i,j}$ .*

*Proof:* We analyze the three cases of Eq. (6).

**Case 1.** Assume that both  $J_{i,j}$  and  $J_{y,z}$  are the first segment of their respective jobs (i.e.,  $j = z = 1$ ). For a starting segment to be able to execute, it needs to be (1) released, (2) the resource it requests (if it requests one) must be available and (3) a core must be available for it to execute on. By definition,  $r_y^{max}$  is an upper bound on (1) and  $SR_{y,z}^{max}(v_p)$  is an upper bound on (2). Regarding (3), we note that because  $J_{y,1}$  and  $J_{i,1}$  are both starting segments, they both compete for the same available cores. Since  $J_{y,1}$  has a higher priority than  $J_{i,j}$ , if both  $J_{y,1}$  and  $J_{i,j}$  are ready and have their shared resources available at the same time, then  $J_{y,1}$  will certainly start before  $J_{i,j}$ . Therefore, only (1) and (2) decide whether  $J_{y,1}$  will start to execute before  $J_{i,j}$ . Thus, if  $J_{i,j}$  did not start before  $\max\{r_y^{max}, SR_{y,z}^{max}(v_p)\}$  when  $z = j = 1$ , then  $J_{y,1}$  will be dispatched before  $J_{i,j}$ .

If  $J_{i,j}$  is not a starting segment (i.e.,  $j > 1$ ), then it already has a reserved core. Therefore, it does not compete with  $J_{y,z}$  for the same core (i.e., we must account for (3)).

**Case 2.** If  $J_{y,z}$  is the first segment of the higher priority job  $J_y$ , then, by definition,  $A_1^{max}(v_p)$  is a safe upper bound on (3). Thus, because  $J_{y,z}$  has a higher priority than  $J_{i,j}$ ,  $J_{y,z}$  will be dispatched before  $J_{i,j}$  if  $J_{i,j}$  did not start to execute before  $\max\{A_1^{max}(v_p), r_y^{max}, SR_{y,z}^{max}(v_p)\}$  when  $z = 1 \wedge j > 1$ .

**Case 3.** If  $J_{y,z}$  is not the first segment of the higher-priority job  $J_y$ , then job  $J_y$  is already released and it already claimed a core. Thus, by definition,  $Cl_y^{max}(v_p)$  is an upper bound on (3). Hence, as  $J_{y,z}$  has a higher priority than  $J_{i,j}$ ,  $J_{y,z}$  will be dispatched before  $J_{i,j}$  if  $J_{i,j}$  did not start to execute before  $\max\{Cl_y^{max}(v_p), SR_{y,z}^{max}(v_p)\}$  when  $z > 1$ . ■

**Lemma 5.** *A segment  $J_{i,j} \in \mathcal{R}^P$  cannot be the direct successor of a state  $v_p$  if it starts executing later than  $t_{high} - 1$ .*

*Proof:* According to Lemma 4,  $\forall J_{y,z} \in \mathcal{H}$ ,  $J_{y,z}$  will start before  $J_{i,j}$  if  $J_{i,j}$  did not start before  $th_{y,z}(J_{i,j})$ . Then,  $J_{i,j}$  will not be the next segment to succeed  $v_p$  if  $J_{i,j}$  did not start before  $\min_{\infty}\{th_{y,z}(J_{i,j}) \mid J_{y,z} \in \mathcal{H}\}$ . ■

It directly follows that an upper bound on the start time of  $J_{i,j}$  can be computed according to Lemma 6.

**Lemma 6.** *A segment  $J_{i,j} \in \mathcal{R}^P$  can be the first segment to start executing in state  $v_p$  only if it starts no later than*

$$LST_{i,j} = \min\{t_{wc}, t_{high} - 1\} \quad (7)$$

<sup>1</sup> $\min_{\infty}\{X\} = \min\{X \cup \{\infty\}\}$  where  $X$  is a set of positive values.

*Proof:* Since  $LST_{i,j} \leq t_{wc}$  (by Cor. 1) and  $LST_{i,j} \leq t_{high} - 1$  (by Lemma 5), the claim holds. ■

### E. Eligibility Condition

Now that we have computed a lower bound  $EST_{i,j}(v_p)$  on the earliest start time and an upper bound  $LST_{i,j}(v_p)$  on the latest start time of  $J_{i,j}$  in state  $v_p$ , it is rather obvious that  $J_{i,j}$  can be a successor to  $v_p$  only if

$$EST_{i,j}(v_p) \leq LST_{i,j}(v_p). \quad (8)$$

**Lemma 7.** *A segment  $J_{i,j}$  is a direct successor of  $v_p$  only if  $EST_{i,j}(v_p) < \infty$  and Inequality (8) holds.*

*Proof:* According to Eq. (2), if  $EST_{i,j}(v_p) = \infty$  then  $J_{i,j}$  is the first segment of job  $J_i$  and no free core is available to start executing  $J_{i,j}$  (i.e., every core is claimed by an unfinished job). Since we assume a non-preemptive system, the scheduler cannot dispatch  $J_{i,j}$  on a core in  $v_p$ . Further, from Lemma 1, we know that  $EST_{i,j}(v_p)$  is a lower bound on the earliest time at which  $J_{i,j}$  can start executing in  $v_p$ . From Lemma 6, we have that  $LST_{i,j}(v_p)$  is an upper bound on the time at which  $J_{i,j}$  can start executing in  $v_p$ . Hence, if  $EST_{i,j} > LST_{i,j}$ , it creates a contradiction, thereby meaning that  $J_{i,j}$  cannot start to execute in  $v_p$  and thus cannot be a successor to  $v_p$ . ■

Beside the eligibility condition stated in Lemma 7, systems using FIFO spin locks have a second eligibility condition.

Consider the case where two segments  $J_{i,j}$  and  $J_{y,z}$  compete for the same shared resource. Since access to the resource is granted in FIFO order, the order in which those job segments will start to execute depends on the order of their requests. Assume that we know a lower bound on the earliest time at which  $J_{i,j}$  may request its shared resource, and that we further have an upper bound on the latest time at which  $J_{y,z}$  may request the resource. We refer to those bounds as  $ERT_{i,j}(v_p)$  and  $LRT_{y,z}(v_p)$ , respectively. We prove in Theorem 1 that if  $ERT_{i,j}(v_p) > LRT_{y,z}(v_p)$ , then  $J_{y,z}$  must execute before  $J_{i,j}$ , thereby implying that  $J_{i,j}$  cannot be the first segment dispatched in state  $v_p$ .

To prove Theorem 1, we first prove a lower and an upper bound on  $ERT_{i,j}(v_p)$  and  $LRT_{y,z}(v_p)$ , respectively.

**Lemma 8.** *Segment  $J_{i,j} \in \mathcal{R}^P$  cannot request a shared resource in  $v_p$  earlier than*

$$ERT_{i,j}(v_p) = \begin{cases} \max\{r_i^{min}, A_1^{min}(v_p)\} & \text{if } j = 1; \\ C_i^{min}(v_p) & \text{if } j > 1. \end{cases} \quad (9)$$

*Proof:* The first segment  $J_{i,1}$  of job  $J_i$  cannot request a resource prior to its release, i.e., not before  $r_i^{min}$ , nor can it request a resource prior to the earliest time at which a core may become available to start its execution, i.e., not before  $A_1^{min}(v_p)$ . Thus, the earliest time at which  $J_{i,1}$  may request its resource is lower-bounded by  $\max\{r_i^{min}, A_1^{min}(v_p)\}$  (hence case 1 of Eq. (9)). If  $J_{i,j}$  is not the first segment of  $J_i$ , then  $J_i$  is already released and  $J_{i,j}$  requests the resource as soon as preceding segment completes, which is lower-bounded by  $C_i^{min}(v_p)$  (hence case 2 of Eq. (9)). ■

**Lemma 9.** *A job segment  $J_{y,z} \in \mathcal{R}^P$  will have certainly requested its shared resource by time*

$$LRT_{y,z}(v_p) = \begin{cases} \max\{r_y^{max}, A_1^{max}(v_p)\} & \text{if } j = 1; \\ C_y^{max}(v_p) & \text{if } j > 1. \end{cases} \quad (10)$$

*Proof:* A segment  $J_{y,z}$  will certainly request its shared resource when (1) the job it belongs to has been released and (2) the segment has a core to execute on.

By definition, the first segment  $J_{y,1}$  of a job  $J_y$  is certainly released at time  $r_y^{max}$  and a core is certainly available by time  $A_1^{max}(v_p)$ . Thus,  $J_{y,1}$  will have certainly requested its resource by time  $\max\{r_y^{max}, A_1^{max}(v_p)\}$  (hence case 1 of Eq. (10)).

If  $J_{y,z}$  is not the first segment of job  $J_y$ , then, because  $J_{y,z} \in \mathcal{R}^P$ , the predecessor of  $J_{y,z}$  must have started to execute. Hence,  $J_y$  is already certainly released, and  $LRT_{y,z}(v_p)$  is only bounded by the time at which the core claimed by the segment of  $J_y$  that precedes  $J_{y,z}$  becomes available to execute  $J_{y,z}$ . That time is upper-bounded by  $C_y^{max}(v_p)$ . Therefore, segment  $J_{y,z}$  will have certainly requested its resource by time  $C_y^{max}(v_p)$  (hence case 2 of Eq. (10)). ■

Now that the *earliest* time  $ERT_{i,j}(v_p)$  at which  $J_{i,j}$  may request its resource, and the *latest* time  $LRT_{y,z}(v_p)$  at which another segment  $J_{y,z}$  has certainly requested its resource have been bounded, we can prove the following necessary condition for  $J_{i,j}$  to possibly be the next segment to start executing in state  $v_p$ .

**Theorem 1.** *A segment  $J_{i,j}$  is a direct successor of  $v_p$  only if  $EST_{i,j}(v_p) < \infty$ , Condition (8) holds, and*

$$\forall J_{y,z} \in \mathcal{R}^P \text{ s.t. } \eta_{y,z} = \eta_{i,j} \neq \emptyset, ERT_{i,j}(v_p) \leq LRT_{y,z}(v_p). \quad (11)$$

*Proof:* The necessity of  $EST_{i,j}(v_p) < \infty$  and Condition (8) were already proven in Lemma 7. Therefore, we focus on proving Condition (11). From Lemma 8, we know that  $ERT_{i,j}(v_p)$  is the earliest time at which a job  $J_{i,j}$  can request its shared resource. Similarly, from Lemma 9, we know that the latest time at which a different segment  $J_{y,z}$  requests its own resource is given by  $LRT_{y,z}(v_p)$ . Therefore, if  $J_{i,j}$  and  $J_{y,z}$  request the same resource  $\ell_x$  and  $ERT_{i,j}(v_p) > LRT_{y,z}(v_p)$ ,  $J_{y,z}$  must certainly be in front of  $J_{i,j}$  in the FIFO queue regulating access to  $\ell_x$ . In this case,  $J_{y,z}$  will certainly execute before  $J_{i,j}$ , so  $J_{i,j}$  cannot be a direct successor of  $v_p$ . ■

### F. Earliest and Latest Finish Times

Since segments execute non-preemptively, if  $J_{i,j}$  is the segment dispatched in system state  $v_p$ , then its earliest finish time ( $EFT_{i,j}(v_p)$ ) and latest finish time ( $LFT_{i,j}(v_p)$ ) are:

$$EFT_{i,j}(v_p) = EST_{i,j}(v_p) + C_{i,j}^{min}(v_p), \quad (12)$$

$$LFT_{i,j}(v_p) = LST_{i,j}(v_p) + C_{i,j}^{max}(v_p). \quad (13)$$

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**Algorithm 2:** Create a new state  $v'_p$  by executing job segment  $J_{i,j}$  after  $v_p$ .

---

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1 if  $J_{i,j}$  is the first segment of  $J_i$  then
2   if  $J_{i,j}$  is not the last segment of  $J_i$  then
3     //  $J_i$  claims a core
4     Create the interval  $Cl_i = [EFT_{i,j}(v_p), LFT_{i,j}(v_p)]$ ;
5      $\mathcal{C}(v'_p) \leftarrow \mathcal{C}(v_p) \cup Cl_i$ ;
6   end
7 else
8   if  $J_{i,j}$  is the last segment of  $J_i$  then
9     // the core claimed by  $J_i$  is released
10     $\mathcal{C}(v'_p) \leftarrow \mathcal{C}(v_p) \setminus Cl_i$ ;
11  else
12    // update the core claimed by  $J_i$ 
13     $Cl_i = [EFT_{i,j}(v_p), LFT_{i,j}(v_p)]$ ;
14     $\mathcal{C}(v'_p) \leftarrow \mathcal{C}(v_p)$ ;
15  end
16 end
17 Update  $\mathcal{C}$  using Eqs. (14) and (15);
18 // Update the shared resource availability
19 if  $\eta_{i,j} \neq \emptyset$  then
20   Let  $\ell_x = \eta_{i,j}$ ;
21    $SF_x^{min}(v'_p) = EST_{i,j}(v_p) + L_{i,j}^{min}$ ;
22    $SF_x^{max}(v'_p) = LST_{i,j}(v_p) + L_{i,j}^{max}$ ;
23 end
24 // Update the free cores availability intervals
25 Initialize PA and CA using Lemma 11 ;
26 Sort PA and CA in non-decreasing order;
27  $\forall 1 \leq x \leq m - |\mathcal{C}(v'_p)|, A_x(v'_p) \leftarrow [PA_x, CA_x]$ ;

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### G. Creating a new State

As discussed in Sec. III-C, we expand the SAG for every segment  $J_{i,j}$  that is eligible according to Theorem 1. For each such segment, we create a new node  $v'_p$  that represents the new state after dispatching  $J_{i,j}$ , as described in Alg. 2.

First, depending on whether  $J_{i,j}$  is the first, last, or an intermediate segment of job  $J_i$ , it will claim a core, release the core it previously claimed, or keep executing on its claimed core, respectively. Therefore, in lines 1–13, Alg. 2 updates the set of claimed cores  $\mathcal{C}$  by either adding, removing, or updating the claimed-core availability interval associated with job  $J_i$ . If a core is claimed or was claimed, then its availability interval is set to the finish time interval  $[EFT_{i,j}, LFT_{i,j}]$  of the segment  $J_{i,j}$  that starts to execute on it (Lines 3 and 10).

For all other claimed cores in  $\mathcal{C}$ , their availability interval is updated according to Equations (14) and (15) below.

**Lemma 10.** *If the first segment executed in system state  $v_p$  starts executing no later than  $EST_{i,j}(v_p)$ , then*

$$Cl_y^{min}(v'_p) = \max\{EST_{i,j}(v_p), Cl_y^{min}(v_p)\} \quad (14)$$

$$Cl_y^{max}(v'_p) = \max\{EST_{i,j}(v_p), Cl_y^{max}(v_p)\} \quad (15)$$

*Proof:* Since the first segment that starts to execute in system state  $v_p$  does so no later than  $EST_{i,j}(v_p)$ , the earlier time at which the next segment may start to execute is not before  $EST_{i,j}(v_p)$ . Thus, the cores are not available to execute new segments before  $EST_{i,j}(v_p)$ . This proves the lemma. ■

Thereafter, the availability interval of the resource accessed by  $J_{i,j}$  (if any) is updated to align with the end of  $J_{i,j}$ 's critical

section, i.e.,  $L_{i,j}^{min}$  time units after it started executing at the earliest, and  $L_{i,j}^{max}$  time units after it started at the latest.

Finally, Alg. 2 uses Lemma 11 to create two sets  $PA$  and  $CA$  that store the times at which each free core becomes possibly and certainly available after  $J_{i,j}$  started to execute (Line 20). Then, as discussed in Sec. III-B, the free cores availability intervals can be computed by sorting the sets  $PA$  and  $CA$  in a non-decreasing order and picking the  $x^{\text{th}}$  element of the sorted set  $PA$  ( $CA$ , resp.) as the lower bound (upper bound, resp.) on the availability interval  $A_x$  (Line 22).

**Lemma 11.** *The times at which each free core becomes possibly and certainly available in  $v'_p$  are contained in the sets  $PA$  and  $CA$ , respectively, computed as follows.*

$$PA = \begin{cases} \max\{EST_{i,j}, A_x^{min}\} \mid 2 \leq x \leq m - |\mathcal{C}| & \text{if } j = 1 \neq n_i \\ \max\{EST_{i,j}, A_x^{min}\} \mid 2 \leq x \leq m - |\mathcal{C}| \cup \{EFT_{i,n_i}\} & \text{if } j = 1 = n_i \\ \max\{EST_{i,j}, A_x^{min}\} \mid 1 \leq x \leq m - |\mathcal{C}| & \text{if } 1 < j < n_i \\ \max\{EST_{i,j}, A_x^{min}\} \mid 1 \leq x \leq m - |\mathcal{C}| \cup \{EFT_{i,n_i}\} & \text{if } j = n_i > 1 \end{cases}$$

$$CA = \begin{cases} \max\{EST_{i,j}, A_x^{max}\} \mid 2 \leq x \leq m - |\mathcal{C}| & \text{if } j = 1 \neq n_i \\ \max\{EST_{i,j}, A_x^{max}\} \mid 2 \leq x \leq m - |\mathcal{C}| \cup \{LFT_{i,n_i}\} & \text{if } j = 1 = n_i \\ \max\{EST_{i,j}, A_x^{max}\} \mid 1 \leq x \leq m - |\mathcal{C}| & \text{if } 1 < j < n_i \\ \max\{EST_{i,j}, A_x^{max}\} \mid 1 \leq x \leq m - |\mathcal{C}| \cup \{LFT_{i,n_i}\} & \text{if } j = n_i > 1 \end{cases}$$

*Proof:* We use the following facts.

**Fact 1.** As already proven for Lemma 10, because  $EST_{i,j}(v_p)$  is the earliest time at which  $J_{i,j}$  starts to execute in system state  $v_p$ , no segment dispatched after  $J_{i,j}$  may start to execute prior to  $EST_{i,j}(v_p)$ . Therefore, no core may be available to execute new jobs before  $EST_{i,j}(v_p)$ . Thus,  $EST_{i,j}(v_p)$  is a lower bound on the availability time of any free core in  $v'_p$ .

**Fact 2.** If  $J_{i,j}$  is not the first segment of job  $J_i$  (i.e.,  $j > 1$ ), then it already has a claimed core and does not execute on a free core. Thus, the availability of each free core in  $v_p$  remains the same in  $v'_p$ . In combination with Fact 1, we get  $PA \supseteq \{\max\{EST_{i,j}(v_p), A_x^{min}(v_p)\} \mid 1 \leq x \leq m - |\mathcal{C}(v_p)|\}$  and  $CA \supseteq \{\max\{EST_{i,j}(v_p), A_x^{max}(v_p)\} \mid 1 \leq x \leq m - |\mathcal{C}(v_p)|\}$ .

**Fact 3.** Because we assume a work-conserving scheduler, if  $J_{i,j}$  is the first segment of  $J_i$  (i.e.,  $j = 1$ ), it will start executing on the first available free core in  $v_p$ . All the other cores will thus keep the same availability interval. Therefore, in combination with Fact 1, we get that  $PA \supseteq \{\max\{EST_{i,j}(v_p), A_x^{min}(v_p)\} \mid 2 \leq x \leq m - |\mathcal{C}(v_p)|\}$  and  $CA \supseteq \{\max\{EST_{i,j}(v_p), A_x^{max}(v_p)\} \mid 2 \leq x \leq m - |\mathcal{C}(v_p)|\}$ .

**Fact 4.** If  $J_{i,j}$  is not the first or last segment of job  $J_i$  (i.e.,  $1 < j < n_i$ ), then it does not release the core claimed by  $J_i$ . Thus, the set of free core in  $v'_p$  is the same as in  $v_p$ .

**Fact 5.** If  $J_{i,j}$  is the last segment of  $J_i$  (i.e.,  $j = n_i$ ), then it releases the core claimed by  $J_i$  at the end of its execution. Thus, the core that was claimed by  $J_i$  in  $v_p$  becomes free for other jobs to execute on in  $v'_p$ . That released core is available



at the earliest at the EFT of  $J_{i,j}$ , and at the latest at the LFT of  $J_{i,j}$ . Hence,  $PA \supseteq EFT_{i,j}(v_p)$  and  $CA \supseteq LFT_{i,j}(v_p)$ .

Fact 3 proves the first case in the definitions of  $PA$  and  $CA$ , respectively. The combination of Facts 3 and 5 prove the respective second cases. Fact 4 proves the third cases, and the combination of Facts 2 and 5 proves the fourth case. ■

#### H. Merge Phase

In order to delay a potential state-space explosion by considering every scenario in a different path, we introduce Rule 1 that merges two nodes of the graph into a single node that covers all states covered by the initial two nodes. This slows down the growth of the SAG (in terms of the number of nodes) while maintaining soundness.

**Rule 1 (Merge rule).** *Two nodes  $v_p$  and  $v_q$  are merged if  $\mathcal{J}^P = \mathcal{J}^Q$  and  $\forall x, 1 \leq x \leq m - |C|, A_x(v_p) \cap A_x(v_q) \neq \emptyset$ .*

When two states  $v_p$  and  $v_q$  are merged into  $v_z$ , each free-core availability interval  $A_x(v_z)$ , claimed core availability interval  $Cl_x(v_z)$ , and shared resource availability interval  $SR_x(v_z)$  in the merged state  $v_z$  are computed so as to fully cover the intervals of the initial states  $v_p$  and  $v_q$ .

$$A_x(v_z) = [\min\{A_x^{min}(v_p), A_x^{min}(v_q)\}, \max\{A_x^{max}(v_p), A_x^{max}(v_q)\}] \quad (16)$$

$$Cl_x(v_z) = [\min\{Cl_x^{min}(v_p), Cl_x^{min}(v_q)\}, \max\{Cl_x^{max}(v_p), Cl_x^{max}(v_q)\}] \quad (17)$$

$$SR_x(v_z) = [\min\{SR_x^{min}(v_p), SR_x^{min}(v_q)\}, \max\{SR_x^{max}(v_p), SR_x^{max}(v_q)\}] \quad (18)$$

We now prove that a merge maintains soundness.

**Lemma 12.** *For two states  $v_p$  and  $v_q$  merged according to Rule 1, the set of claimed cores in  $v_p$  and  $v_q$  are assigned to the same job segments.*

*Proof:* Every job that has started and did not yet finish its execution until reaching state  $v_p$  has a claimed core in  $v_p$ . Since by Rule 1, the set of job segments in the path to  $v_p$  and  $v_q$  are identical (i.e.,  $\mathcal{J}^P = \mathcal{J}^Q$ ), all jobs that claimed core in  $v_p$  must also have a claimed core in  $v_q$ , and all jobs that claimed a core in  $v_q$  must also have claimed a core in  $v_p$ . ■

**Lemma 13.** *For two states  $v_p$  and  $v_q$  merged according to Rule 1, the number of free-core availability intervals are the same in  $v_p$  and  $v_q$ .*

*Proof:* Lemma 12 proves that the claimed cores in  $v_p$  and  $v_q$  are assigned to the same job segments. Thus, the number of claimed cores must be the same in  $v_p$  and  $v_q$ . Hence, the number of free cores must also be the same since the sum of claimed and free cores is always equal to  $m$ . It follows that we have as many free-core availability intervals in  $v_p$  as in  $v_q$ . ■

**Theorem 2.** *Merging two states  $v_p$  and  $v_q$  according to Rule 1 and Equations 16, 17, and 18 is safe, i.e., it does not remove any potentially reachable system state from the graph.*

*Proof:* Rule 1 ensures that the sets of segments that have started executing on the path to  $v_p$  and  $v_q$  are identical for  $v_p$  and  $v_q$ . Hence, the set of segments that still need to execute in the merged state  $v_z$  is the same as in  $v_p$  and  $v_q$ . According to Lemmas 12 and 13, the set of claimed cores and the number of free cores in state  $v_p$  and  $v_q$  are the same, and are thus the same in the merged state too. Since the availability interval of shared resources and free and claimed cores in the merged state  $v_z$  are the union of those in  $v_p$  and  $v_q$ , any possible combination of a given number of cores and set of resources becoming available at a given time that is possible in either state  $v_p$  or  $v_q$  is also possible in the merged state  $v_z$ . Hence, all sequences of segment executions that may follow from  $v_p$  and  $v_q$  are also possible in  $v_z$  and the set of all system states reachable from  $v_z$  includes every state that is reachable from the original states  $v_p$  and  $v_q$ . ■

#### I. Correctness

We now put all the pieces together and establish soundness.

**Theorem 3.** *For any execution scenario such that segment  $J_{i,j}$  completes at  $t$ , there is a path  $\langle v_1, \dots, v_p, v'_p \rangle$  in the SAG such that  $J_{i,j}$  is the label of the edge connecting  $v_p$  to  $v'_p$  and  $t \in [EFT_{i,j}(v_p), LFT_{i,j}(v_p)]$ .*

*Proof:* Assume that there is a path  $\langle v_1, \dots, v_p \rangle$  such that the claim is respected for all segments that started to execute before  $J_{i,j}$  in the execution scenario that led  $J_{i,j}$  to finish at time  $t$ . Furthermore, assume that the availability intervals of state  $v_p$  correctly bound the actual availability times of the shared resources, free cores, and claimed cores.

We prove that  $t \in [EFT_{i,j}(v_p), LFT_{i,j}(v_p)]$ , that a new system state  $v'_p$  is created from scheduling  $J_{i,j}$  in  $v_p$ , and that the availability intervals of state  $v'_p$  correctly bound the actual availability times of the shared resources, free cores, and claimed cores after executing  $J_{i,j}$ .

Under the inductive assumption stated above, Lemma 1 and Lemma 6 prove that  $EST_{i,j}(v_p)$  and  $LST_{i,j}(v_p)$  are safe lower- and upper-bounds on the start time of  $J_{i,j}$ , respectively. Because segments execute non-preemptively, Equations (12) and (13) are thus safe lower- and upper-bounds on  $t$  (i.e., we proved that  $t \in [EFT_{i,j}(v_p), LFT_{i,j}(v_p)]$ ). Further, by the inductive assumption, the condition of Theorem 1 must hold for  $J_{i,j}$  and Line 7 of Alg. 1 ensures that the graph is expanded with a new node  $v'_p$ . Then, by Lemmas 10 and 11 and the discussion of Alg. 2, the availability intervals of  $v'_p$  correctly bound the actual availability of shared resources and cores after executing  $J_{i,j}$ . Therefore, the inductive assumption is respected for  $v'_p$ . Finally, according to Lemma 10, potentially merging  $v'_p$  with another node (lines 14 to 18 of Alg. 1) maintains the validity of the inductive assumption.

Crucially, the inductive assumption (i.e., correct availability intervals) obviously holds for  $v_1$  (in which all cores and shared

resources are supposed to be available) and thus follows by induction on all the states created by Alg. 1. ■

## V. EMPIRICAL EVALUATION

We conducted a large-scale schedulability study to evaluate any gains in accuracy and to test for scalability limitations.

### A. Setup and Workloads

To obtain a large corpus of diverse workloads with varied contention characteristics, we generated task sets as follows. For a given number of cores  $m \in \{2, 4\}$ , we generated  $n \in \{m + 1, 2m, 3m, 5m\}$  periodic tasks that shared  $n^r \in \{2m, 5m\}$  resources. Each task was configured to have a number of critical sections chosen uniformly at random from  $\{0, 1, \dots, n^{cs}\}$ , for  $n^{cs} \in \{5, 15\}$ . The length of each critical section was drawn uniformly at random from either  $[1\mu s, 15\mu s]$ ,  $[10\mu s, 50\mu s]$ , or  $[50\mu s, 150\mu s]$  (*short*, *intermediate*, or *long* critical sections, respectively). Additionally, to obtain Fig. 2b, task sets for a specific setup with  $m = 6$ ,  $n = 12$ ,  $n^{cs} = 5$  and  $n^r = 12$  were generated as well.

For each considered combination of  $m$ ,  $n$ ,  $n^r$ ,  $n^{cs}$ , and critical section lengths, we varied the total utilization  $U$  from 5% to 95% in steps of 5, and for each  $U$ , we generated 250 task sets, for a total of more than 450,000 task sets.

For a given  $U$  and  $n$ , we generated  $n$  per-task utilizations  $u_1, u_2, u_3, \dots$  that sum to  $U$  using Emberson et al.’s *RandFixed-Sum* method [17]. To reflect that nonpreemptive scheduling is used in practice only for workloads with short jobs, for each task, an initial cost value  $C'_i$  was drawn from a normal distribution with mean  $3ms$  and standard deviation  $1.5ms$  (values of less than  $10\mu s$  were re-drawn). To avoid unrealistic periods, we then selected the task period  $T_i \in \{5, 8, 10, 12, 15, 20, 25, 30, 50, 75, 100, 150, 200, 250, 300, 500\}ms$  closest to  $C'_i/u_i$  and set the task’s WCET to  $u_i \cdot T_i$  (but no less than the sum of maximum critical section lengths).

The total WCET was randomly distributed across the task’s segments  $C_{i,1}^{max}, C_{i,2}^{max}, \dots$  while respecting the corresponding maximum critical-section lengths (if any). Each segment was assigned a BCET  $C_{i,j}^{min}$  by drawing a value uniformly at random from  $[0.1C_{i,j}^{max}, 0.5C_{i,j}^{max}]$ . Each critical section was assigned a minimum length of  $0\mu s$  with probability 0.25, and otherwise chosen at random as  $L_{i,j}^{min} \in [0.1L_{i,j}^{max}, 0.5L_{i,j}^{max}]$ .

Finally, tasks were assigned unique priorities using the DkC heuristic for fixed-priority scheduling [14]. Afterwards, as a necessary test, one hyperperiod of the task set was simulated assuming that each job executes for its WCET (without any blocking). If the simulation already exhibited a deadline miss, then the task set was discarded to avoid generating task sets that are obviously infeasible under nonpreemptive scheduling.

### B. Implementation and Baselines

We implemented the proposed analysis, which we denote as **{EDF, FP}-SAG-SR** in the following (under global nonpreemptive EDF and FP scheduling, respectively), by extending Nasri et al.’s open-source SAG tool [2,23].

As there exists no directly comparable analysis to compare against (the proposed solution is the first of its kind), we

further considered the following loosely related baselines to provide context: 1) **{EDF, FP}-SAG-NO-BLOCKING**: Nasri et al.’s analysis [23] *without any blocking*. This analysis is not sound in the presence of shared resources; it merely serves to indicate an upper bound on attainable schedulability. 2) **{EDF, FP}-SAG-INF**: Nasri et al.’s analysis [23] with an *inflation-based blocking analysis* where each job’s WCET is increased prior to response-time analysis to account for possible blocking based on the holistic blocking analysis approach [9, Ch. 5.4]. This analysis is sound, but structurally pessimistic (not scenario-aware), and thus provides a simple lower bound on schedulability that the proposed analysis should exceed. 3) **EDF-NO-BLOCKING**: the schedulability tests for *preemptive* global EDF provided in the open-source SchedCAT library [1] *without any blocking*, as an upper bound on schedulability under preemptive EDF scheduling. 4) **FP-NO-BLOCKING**: similarly, the schedulability tests for *preemptive* global FP scheduling in SchedCAT without any blocking. 5) **{EDF, FP}-FMLP-SHORT**: inflation-based holistic blocking analysis [9] of non-preemptive FIFO-ordered spin locks (i.e., “FMLP for short resources” [7]), as provided in SchedCAT. 6) **{EDF, FP}-OMLP**: inflation-based holistic blocking analysis of the *suspension-based* global OMLP locking protocol [11], as provided in SchedCAT. 7) **FP-FMLP-LONG**: Yang et al.’s analysis [28] of the suspension-based “FMLP for long resources” [7]. 8) **FP-PIP**: similarly, Yang et al.’s analysis [28] of the suspension-based PIP [16,25].

### C. Results

In total, our experimental setup considered 14 individual analyses for over 95 different scenarios (i.e., parameter combinations). Due to space constraints, we cannot report the complete set of results here; a representative selection of our results is shown in Fig. 2 with the specific parameter choices given in each plot. To avoid clutter, only a subset of the 14 curves is shown in each graph.

Fig. 2a shows a comparison of the various SAG analyses for  $m = 4$ . First, there is little difference w.r.t whether jobs are prioritized according to fixed task-level DkC priorities [14] or by job-level EDF priorities. More importantly, however, the results show a large gain in schedulability relative to the inflation-based baseline. For example, while FP-SAG-INF only deemed about 15% of the task sets to be schedulable at 45% total utilization, FP-SAG-SR shows that more than 60% of the tested task sets are actually schedulable. Overall, the proposed analysis closes *more than half of the gap* between the inflation-based baseline and the upper bound on attainable schedulability represented by **{EDF, FP}-SAG-NO-BLOCKING**, which highlights a significant reduction in pessimism.

Fig. 2b focuses on EDF scheduling and shows similar trends for  $m = 6$ , with EDF-SAG-SR attaining much higher schedulability than EDF-SAG-INF. Furthermore, a comparison with EDF-OMLP and EDF-FMLP-SHORT shows the proposed analysis to be competitive with the state of the art for *preemptive* systems (for the considered workloads), especially considering that the EDF-NO-BLOCKING baseline

reveals that preemptive scheduling has a slight advantage in this scenario. Note that our evaluation does not reflect any differences in scheduling and runtime overheads, which can be expected to be (much) lower under nonpreemptive scheduling. Fig. 2c shows largely identical trends for  $m = 4$ .

Fig. 2d shows an analogous comparison for FP scheduling; here preemptive scheduling has a significant advantage and Yang et al.’s suspension-aware analysis [28] of the PIP and the “long” FMLP variant stand out as particularly effective. However, note again that this is a somewhat lopsided comparison due to workload differences (preemptive vs. nonpreemptive) and since we are discounting lower spin-lock overheads.

Fig. 2e shows the effect of increasing the number of tasks. Generally, schedulability drops as  $n$  increases, which is unsurprising as an increase in tasks is correlated with an increase in contention. Nonetheless, significant accuracy gains are apparent compared to the inflation-based baseline. For example, for  $n = 5$  and  $U = 60\%$ , FP-SAG-SR shows more than 50% of the workloads to be schedulable, whereas FP-SAG-INF can show this for less than 20% of the workloads.

In Fig. 2f, we look at the effect of varying the number of cores. Here we clearly see that, for all investigated values of  $m$ , the proposed analysis performs better than the inflation-based baseline. Schedulability drops with increasing  $m$  since an increase in parallelism corresponds to an increase in contention (since  $n = 2m$  in the shown scenarios).

In Figs. 2g and 2h, we investigate the runtime of the algorithm as a function of the total utilization, number of tasks and number of cores. Since the experiments have been run on several different machines (all being server-class machines with large amounts of DRAM), we cannot relate the runtime results to a single machine configuration. However, there are still interesting trends to be observed. For example, in Fig. 2g, the runtime increases with the utilization until  $U \approx 30\%$ . From 30% onwards, however, the runtime decreases. This is explained by the fact that when the utilization increases, we deal with increased contention, which results in more branching during the analysis, thereby leading to an increased runtime. However, when the utilization becomes larger, there is also an increasing chance to find more task sets to be unschedulable at an early stage of the analysis. Therefore, the graph construction is stopped earlier and the runtime decreases. We also observe in Fig. 2g that the runtime increases significantly with the number of tasks, which we link to an increasing number of jobs in the observation window and thus an increasing number of possible job segments execution orderings. In Fig. 2h we investigate the impact of varying the number of cores on the runtime (note that the number of tasks also changes in that experiment since  $n = 2m$ ). Non-surprisingly, we observe that the runtime increases exponentially with the number of tasks and cores. However, even in the worst configuration shown in Fig. 2h, 95% of the generated task sets have an analysis runtime below 2.7 min. In the configurations of Fig. 2g, more than 95% of the task set are analyzed in less than 40 min. when  $n \leq 12$ , but task sets start to time out when  $n = 20$ .

Our evaluation also revealed some limitations. First, note

how in Fig. 2e the results for FP-SAG-SR for  $n = 20$  tasks are worse than for the baseline FP-SAG-INF, but *not* for lower task counts. This is a result of the increased runtime of the proposed analysis. In our experiments, we configured a timeout of 120 minutes per task set. As the runtime of the analysis increases with  $n$ , FP-SAG-SR started to exhibit a significant number of timeouts for  $n = 20$ . Scalability limitations also meant that while it is certainly possible to analyze a given system and even perform experiments for a chosen range of parameters when  $m > 4$  (e.g., Fig. 2b shows results for  $m = 6$ ), the general increase in runtime made an evaluation across the entire parameter space (i.e., *hundreds of thousands of workloads*, recall Sec. V-A) impractical for  $m \geq 6$ .

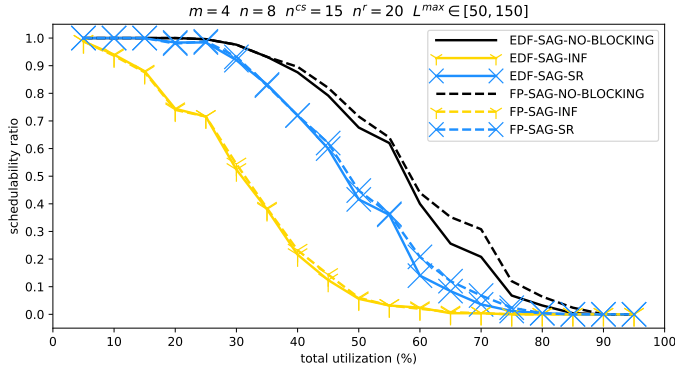
Surprisingly, we also found some rare cases where {EDF, FP}-SAG-SR perform (slightly) worse than the corresponding {EDF, FP}-SAG-INF analyses even in the absence of timeouts. Further investigation led us to discover a source of pessimism related to the encoding of shared-resource availability. Our current state abstraction assumes that the resource remains locked (and thus unavailable) for the whole time from  $SR^{min}$  until  $SR^{max}$  in the worst case. Although accurate in most cases, this interval may become very long in comparison to the actual critical section length in corner cases. There are thus rare situations where even an inflation based test may perform better. Such occasional instances can be easily handled by running both analyses and retaining the better result.

Nonetheless, we overall conclude that *almost* always the proposed analysis shows substantial accuracy gains in comparison to the state of the art and, for the considered workloads, is competitive with (or even superior to) solutions designed for preemptive systems. The positive results virtually across the board provide ample motivation to further improve accuracy and scalability of the proposed approach in future work.

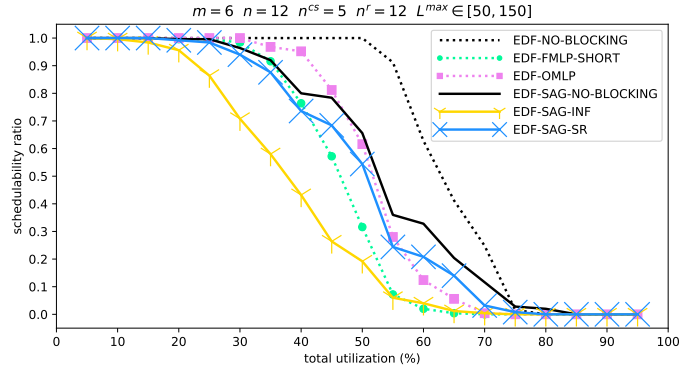
## VI. CONCLUSION

We presented a new WCRT analysis for global JLFP scheduling and non-preemptive tasks with repeating job-release patterns that share resources protected by FIFO spin locks. To the best of our knowledge, it is the first solution to this problem. Our analysis implicitly explores all possible execution and resource access orderings using a novel system-state abstraction that explicitly models resource contention.

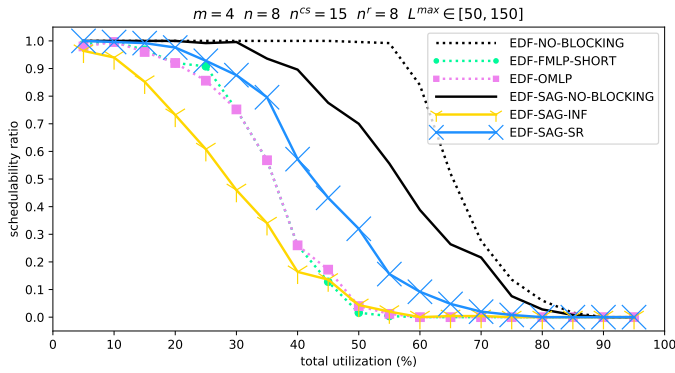
A comparison with inflation-based analyses has shown that our work is substantially less pessimistic. The empirical evaluation has also shown that, for the type of workloads considered in the experiments, our solution comes much closer to hypothetical blocking-free upper bounds and is competitive even with solutions for preemptive systems, suggesting that our analysis successfully discards many more scenarios that reflect impossible combinations of shared resources access orders, low-priority blocking and/or high-priority interference. We also identified an intermittent source of pessimism, which we plan on addressing at the same time as we work on scaling the analysis to larger numbers of core counts using more aggressive pruning techniques (such as partial-order reduction) to disregard scenarios that do not contribute to the worst case.



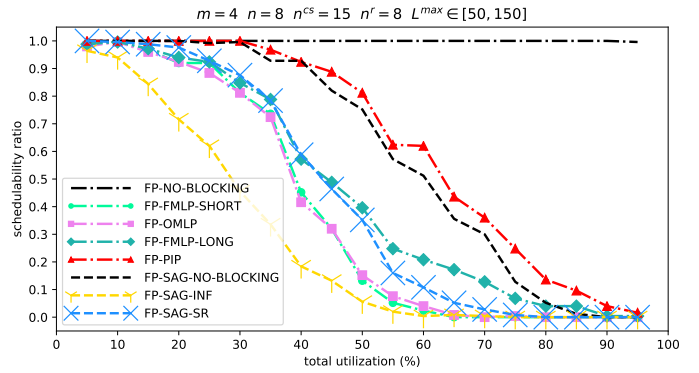
(a) comparison of SAG analyses for  $m = 4$



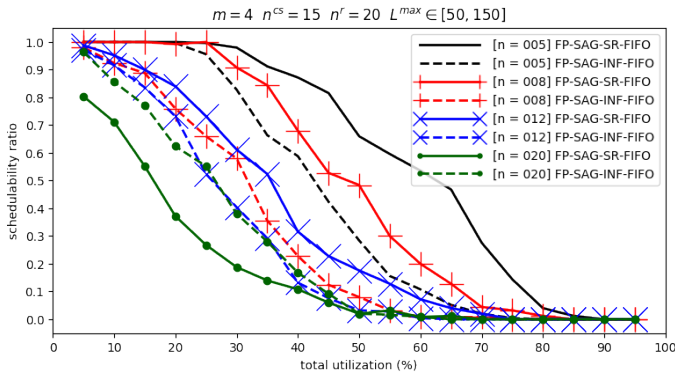
(b) comparison of EDF analyses for  $m = 6$



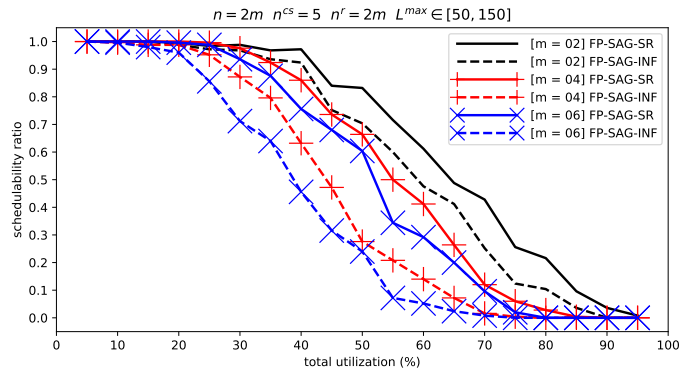
(c) comparison of EDF analyses for  $m = 4$



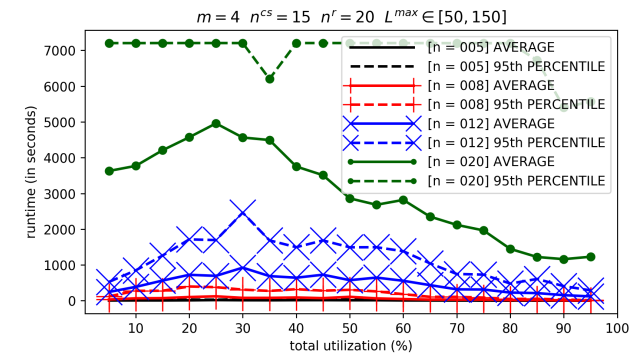
(d) comparison of FP analyses for  $m = 4$



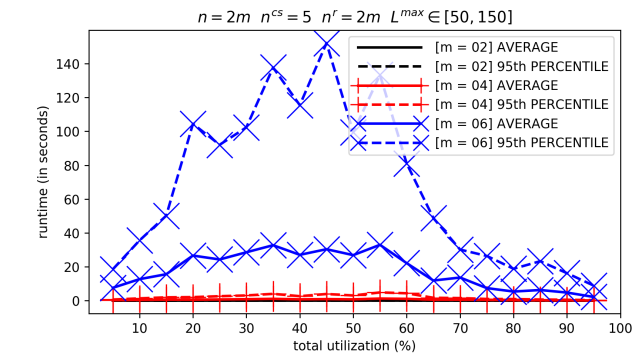
(e) schedulability for varying task set size  $n$



(f) schedulability for varying processor count  $m$



(g) FP-SAG-SR runtime for varying task set size  $n$



(h) FP-SAG-SR runtime for varying processor count  $m$

Fig. 2: Select experimental results. See Sec. V-B for an explanation of the considered analyses.

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