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Speed scaling with power down scheduling for agreeable deadlines[☆]

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ABSTRACT

We consider the problem of scheduling on a single processor a given set of n jobs. Each job j has a workload w_j and a release time r_j. The processor can vary its speed and hibernate to reduce energy consumption. In a schedule minimizing overall consumed energy, it might be that some jobs complete arbitrarily far from their release time. So in order to guarantee some quality of service, we would like to impose a deadline $d_j = r_j + F$ for every job j, where F is a guarantee on the *flow time*. We provide an $O(n^3)$ algorithm for the more general case of *agreeable deadlines*, where jobs have release times and deadlines and can be ordered such that for every $i < j$, both $r_i \le r_j$ and $d_i \le d_j$.

1. Introduction

Recent research addresses the issue of reducing the amount of energy consumed by computer systems while maintaining satisfactory level of performance. This can be done at different levels of a computer system. One possibility is to specify a good scheduling mechanism in the operating system level. Here we have two mechanisms at hand. One common method for saving energy is the *power-down mechanism*, which is to simply suspend the system during long enough idle times. Another common method is *speed scaling*, which is to adjust the processor speed low enough to meet the jobs requirements. In this paper we study the problem of designing scheduling algorithms for minimizing the consumed energy using both mechanisms.

The question whether this problem can be solved in polynomial time was posed by Irani and Pruhs [1], who called it *speed scaling with power down scheduling* problem. We provide an $O(n^3)$ algorithm in this paper for the special case of agreeable deadlines. Jobs may be released at different time moments, and may have distinct deadlines. The agreeable deadline property just means that later released jobs also have later deadlines. This holds, for example, when the deadline of each job is exactly F units after its release time, which arises when one wants to maintain a guarantee of service for the flow time of the jobs.

2. Problem definition

An instance of our scheduling problem consists of n jobs, $1, 2, \ldots, n$, where each job j, $1 \le j \le n$, is specified by a release time/deadline interval[1] $[r_j, d_j)$ in which it must be scheduled and a workload w_j. An instance has the *agreeable deadlines property* if the jobs can be renumbered such that both their release times and deadlines are in non-decreasing order, i.e. $i < j$ implies $r_i \leqslant r_j$ and $d_i \leqslant d_j$.

A schedule is defined by three functions

$$\begin{aligned} \text{mode} &: \mathbb{R} \to \{\text{on}, \text{off}\} \\ \text{speed} &: \mathbb{R} \to \mathbb{R}^+ \\ \text{job} &: \mathbb{R} \to \{\text{none}, 1, \ldots, n\}, \end{aligned}$$

with the following properties

1. $\forall t: \text{speed}(t) > 0 \Rightarrow \text{mode}(t) = \text{on}$
2. $\forall t: \text{speed}(t) = 0 \Leftrightarrow \text{job}(t) = \text{none}$
3. $\forall t: \text{job}(t) = j, j \neq \text{none} \Rightarrow t \in [r_j, d_j)$
4. $\forall j \neq \text{none}: \int \text{speed}(t)dt = w_j$ where the integral is over all times t such that $\text{job}(t) = j$
5. for every time t, there is a positive length interval $I \ni t$ on which the schedule is constant. Moreover I is of the form $(-\infty, u)$, $[t', u)$ or $[t', +\infty)$ for some time points t', u.

The last property is in fact a simplifying assumption to avoid degenerate schedules. The interpretation is that at a time t where $\text{job}(t) = \text{none}$, the machine is idle but switched on when

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[1] Notation: $[t_0, t_1)$ stands for the half open interval $\{t : t_0 \le t < t_1\}$.

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mode(t) = on and is shut down when mode(t) = off. There is a non-negligible energy consumption during the idle time periods, but one avoids the cost of shutting down and rebooting the machine.

The cost (i.e., the consumed energy) of a schedule is specified by three parameters: an exponent $\alpha \in [2, 3]$, a wake-up cost $L > 0$ and a ground dissipation energy $g > 0$, and it has two components:

1. The *speed cost*, that is the energy consumed in all times t such that job(t) = $j \neq$ none. This cost is defined as $c_{\text{speed}} = \int \text{speed}(t)^\alpha \text{d}t$.
2. The *mode cost*, that is the cost of the ground dissipation energy plus the wake-up energy.

A schedule with property (5) partitions the time into a sequence S of disjoint, inclusion-wise maximal intervals, such that mode(t) = on if and only if $t \in \cup S = \cup_{I_k \in S} I$. The sequence S is called the *support* of the schedule, and the energy consumption generated by this support constitutes its mode cost which is defined as $c_{\text{mode}} = L(|S| + 1) + g|\cup S|$. Note that we count a wake-up cost L for the two half-infinite intervals surrounding S.

Hence, the total cost is just the sum $c_{\text{speed}} + c_{\text{mode}}$ and the problem studied in this paper consists in finding a minimum cost schedule for an agreeable deadline instance.

The outline of the paper is as follows. After a brief survey over related work, we start showing structural properties of optimal schedules, in particular we introduce the notion of prefix and suffix of a block. Finally we define the dynamic program, prove its optimality and analyze its complexity, while mentioning implementation issues. The algorithm has been implemented in Python and can handle instances of 300 jobs within a second.

3. Previous work

Our scheduling problem for general instances (with non-agreeable deadlines) was raised in Ref. [1]. No polynomial time algorithm is known for this problem, nor has it been shown to be NP-hard. The current best positive result is a 137/117-approximation provided in [2]. The same paper also gives an NP-hardness proof, however for a different energy mode, when c_{speed} is defined as $\int f(\text{speed}(t))\, \text{d}t$ for some piecewise linear monotone function f.

The general problem contains two subproblems, which have been studied and solved individually. The first one does not consider speed scaling, and restricts to speeds 0 or 1, depending on the mode. Here essentially the goal is to minimize c_{mode} only. This subproblem has been solved in $O(n^5)$ time by dynamic programming [3]. For agreeable instances the complexity has been improved to $O(n^2)$ [4]. The second subproblem does not consider the power down mechanism, and restricts to the single mode 'on' and to ground dissipation energy $g = 0$. Here the problem is to minimize c_{speed} only. This problem has been solved by a widely celebrated greedy algorithm due to Yao et al. [5] in $O(n^3)$ time. The complexity of this algorithm, known as YDS, has been improved to $O(n^2 \log n)$ in Ref. [6] and even to $O(n^2)$ for agreeable instances [7].

Different variants of this problem have been studied in the past years, which include the online setting, as well as different objective values, like minimizing throughput or flowtime. We refer to [8] for an overview.

An important ingredient for the algorithm presented in this work, is the aforementioned YDS algorithm. For completeness we roughly sketch it now. At each step of the algorithm, the interval $I^\star$ of maximum density is selected among all $O(n^2)$ intervals of the form $[r_i, d_j]$. The *density of an interval* is defined as the ratio $s = W/(d_j - r_i)$, where W denotes the total workload of all jobs k with $[r_k, d_k] \subseteq [r_i, d_j]$. The key idea is that any feasible schedule must have for $[r_i, d_j]$ an average speed at least s. Then all those jobs are scheduled in $I^\star$ at speed s using the EDF (Earliest-Deadline-First) policy. No job will miss its deadline by maximal density of $I^\star$. For the sequel of the algorithm the time interval $I^\star$ is *blacked out*. This means that when computing densities of candidate intervals for subsequent iterations, the blacked out intervals are excluded, and the schedule for the remaining jobs must exclude them as well. The algorithm ends when all jobs are scheduled.

4. Structure of an optimal schedule

In this section we show some structural properties of optimal schedules, see Fig. 1 for illustration.

When a job j is running at speed s its execution takes w_j/s time units and the consumed energy is $(s^\alpha + g)w_j/s$. This amount of energy is minimum for speed $s^\star := (g/(\alpha - 1))^{1/\alpha}$, which we call the *critical speed*. Note that $s^\star$ is job independent. The *density* of an interval I is defined as $\sum w_j/|I|$ over all jobs j with $[r_j, d_j) \subseteq I$. An interval is called *dense*, if its density is at least $s^\star$, and *sparse*, otherwise.

Lemma 1 ([9]). *Given an instance of the speed scaling with power down scheduling problem, there is an optimal schedule* (mode, speed, job) *with the following properties.*

job span *for every time t, if* job(t) = $j \neq$ none *then for all times* $u \in [r_j, d_j)$ *with* mode(u) = on, *we have* speed(u) $\geqslant$ speed(t)
earliest deadline first *for every time pair* $t < u$ *if* job(t) $\neq$ none $\neq$ job(u), *then* job(t) $\leqslant$ job(u).
dense intervals *dense intervals I are scheduled according to the YDS rule.*
domination *for any other optimal schedule* (mode′, speed′, job′), *and a smallest time t such that* mode(t) $\neq$ mode′(t) *we have* mode(t) = on *and* mode′(t) = off.

In particular the first property implies that whenever j is scheduled, the speed is the same. The next two properties imply that dense intervals divide the problem into independent subproblems, as we describe now.

Definition 1. *A subinstance of our problem is specified by a pair* (i, j) *with* $i \in \{1, \ldots, n\}$, $j \in \{i - 1, \ldots, n\}$. *For convenience we denote* $d_0 = r_1 - L/g$ *and* $r_{n+1} = d_n + L/g$. *It consists of the interval* $I = [d_{i-1}, r_{j+1})$ *and a job set J. If* $i = j + 1$, *then* $J = \emptyset$, *else* $J = \{i, \ldots, j\}$. *The release time/deadline intervals of these jobs are restricted by intersection to I.*

Note that in case $d_{i-1} < r_{j+1}$ or $d_{i-1} < d_i$ or $r_j < r_{j+1}$, the subinstance (i, j) is infeasible as the release time/deadline interval of i or j is restricted to the empty interval.

We extend also the definition of the cost function for subinstances. The schedule of a subinstance (i, j), consisting of job set J and interval I, is defined by the functions speed : $I \to \mathbb{R}^+$, mode : $I \to$ {on, off} and job : $I \to \{\text{none}\} \cup J$. For the mode cost, let $S := \{t \in I : \text{mode}(t) = \text{on}\}$ be the support of the schedule, and k be the number of intervals in $I \setminus \cup S$. Then, $c_{\text{mode}} := kL + g|\cup S|$. The interpretation is that if immediately before and after I the machine is on, then shutdown intervals at the borders of I also do generate a wake-up cost.

We choose d_0 far enough from r_1 such that w.l.o.g. an optimal schedule for the subinstance $(1, k)$ will start with a shutdown interval. A symmetric property is true for subinstances of the form (k, n). Therefore the cost of the subinstance $(1, n)$ is consistent with the cost definition for the complete instance. Note that the optimum for a subinstance of the form $(i, i - 1)$ equals $\min\{L, g(r_i - d_{i-1})\}$.

Now consider all inclusion-wise maximal dense intervals. They partition the time line into a sequence of alternating dense and sparse intervals.

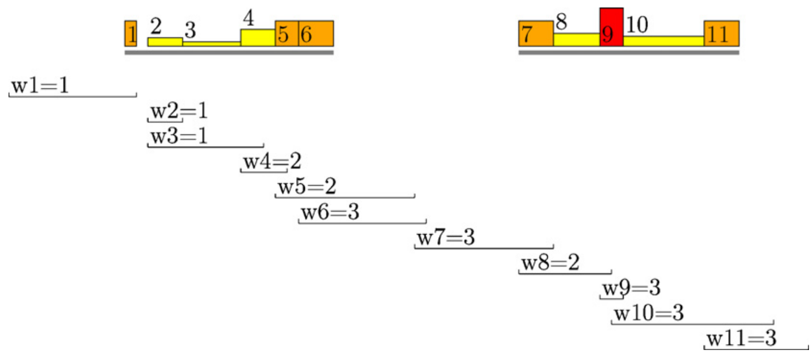

Fig. 1. Structure of an optimal schedule for an instance of 11 jobs. The boxes represent job executions where the height equals speed and the area equals the workload of the job. The different colors of the boxes distinguish critical speed, less than critical speed and more than critical speed. The thick line below represents the mode. Finally the intervals at the bottom represent the release time/deadline intervals of each job j, labeled by its workload w_j. Here the schedule consists of 2 blocks separated by a shutdown interval. Jobs 1,5,6,7,11 are scheduled at critical speed $s^\star$, while job 9 is scheduled with higher speed, as $[r_9, d_9)$ is a dense interval. Note that the schedule is idle but not shutdown between jobs 1 and 2.

The following lemma follows directly from the definitions. We stress here that independence of the subschedules is implied by the agreeable deadline assumption. Lemma 1 states that there is an optimal schedule, satisfying the *earliest deadline first property*, which means that whenever job j is scheduled, all jobs $i<j$ already completed. So the agreeable deadline assumption happens to be quite strong, which permits a dynamic programming approach. However the problem does not become trivial, since one still needs to decide when the machine is to be shutdown and when to be idle.

Lemma 2. *Sparse intervals I are associated to pairs (i, j), such that the portion of an optimal schedule for the original instance restricted to I, is also an optimal schedule for the subinstance (i, j). Moreover none of these subinstances contain dense intervals.*

5. Suffixes and prefixes

In this section, we consider an optimal schedule of an arbitrary subinstance consisting of a job set J and an interval I such that all subintervals of I are sparse. Whenever, in this section, we refer to release times/deadlines r_k, d_ℓ, they are restricted to I (see Fig. 2).

Lemma 3. *For all times $t \in I$,* speed$(t) \leqslant s^\star$.

Proof. Let t be a time that maximizes speed(t), and assume speed$(t) > s^\star$ for the sake of contradiction. We consider an inclusion-wise maximal interval $A \ni t$ on which the speed is constantly speed(t). Let $i, \ldots, k$, $i \leqslant \text{job}(t) \leqslant k$, be the jobs scheduled in this interval. If $A = [r_i, d_k)$, then A is a dense interval, a contradiction to Lemma 2. Thus, the inclusion $A \subseteq [r_i, d_k)$ is strict. Assume $d_k > u$ for $u = \max A$ (the other case is symmetric). By Lemma 1, we have mode(u) = off, and there is a time t' such that job k is scheduled in $[t', u)$. For a small enough $\delta \geqslant 1$, the execution of job k can be extended to $[t', u')$ for $u' = t' + \delta(u - t')$ and lower its speed to speed$(t)/\delta$. This strictly decreases the overall cost, a contradiction to the optimality of the schedule. □

The support of the schedule consists of blocks separated by shutdown intervals. We shall show now that the boundaries of these blocks have a particular structure (see Fig. 2).

Definition 2. *A suffix is a job pair (a, b) such that all jobs $a, \ldots, b$ are scheduled at critical speed between r_a and u with $u = r_a + (w_a + \ldots + w_b)/s^\star$, and* mode$(u)$ = off. *The definition of a prefix is just symmetric.*

Lemma 4. *Let $[t, u)$ be an inclusion-wise maximal shutdown interval in I, that is* mode(t') = off *for all $t' \in [t, u)$. If t is not the start of I, then there is a suffix (a, b) ending at $t = r_a + (w_a + \ldots + w_b)/s^\star$. If u is not the end of I, then there is a prefix $(b+1, c)$ starting at $u = d_c - (w_{b+1} + \ldots + w_c)/s^\star$. Moreover, if both cases hold ($\inf I < t < u < \sup I$) then without loss of generality $r_{b+1} > t$.*

Proof. Suppose that there is an execution interval $[t_0, t)$ where some job b = job(t_0) is scheduled at speed$(t_0) < s^\star$. For a small enough $\delta > 1$ let $t' := t_0 + (t - t_0)/\delta$. Consider a new schedule where the execution interval is *compressed* to $[t_0, t')$, the speed in there is multiplied δ, and the shutdown interval is extended to $[t', u)$. This new schedule has a strictly decreased cost, contradicting optimality.

This shows that if t is not the start of I, then some job b is scheduled right before t, say in some interval $[t_0, t)$, at critical speed. We will now show that there is a job a such that between r_a and t, jobs $a, \ldots, b$ are all scheduled at critical speed. If $t_0 = r_b$, we simply set $a = b$. Otherwise assume that $r_b < t_0$. If right before t_0 the schedule mode is off, then we can slightly shift the execution interval of b to $[t_0 - \varepsilon, t - \varepsilon)$, to obtain a schedule of the same cost but with dominating work toward the beginning. W.l.o.g. we can assume that right before t_0 a job $b-1$ is scheduled in some interval $[t_1, t_0)$. By the job span property of Lemma 1 and Lemma 3 it is scheduled at speed $s^\star$. We iterate the arguments on t_1 and $b-1$, eventually reaching a job a with the required property.

The same argument applied symmetrically shows the existence of a prefix $(b+1, c)$ if u is not the end of I. Now if both suffix and prefix exist, and $r_{b+1} \leqslant t$, then we could shift the execution of $b+1$ from $[u, w_{b+1}/s^\star)$ to $[t, w_{b+1}/s^\star)$, yielding a schedule with more work dominating toward the beginning. The cost of the new schedule remains either the same or it is reduced by L, if $b+1$ were alone in its block. Therefore we can assume w.l.o.g. that $t < r_{b+1}$. □

To proceed to our dynamic programming algorithm we need one more property of suffixes and prefixes implied by the following definition.

Definition 3. *For a given subinstance (i, j) we define two functions $f, h : \{i, \ldots, j\} \to \{i, \ldots, j\}$ as follows: $f(a)$ is the highest index job $b \leq j$ such that for all $a \leq k < b$, $r_a + (w_a + \ldots + w_k)/s^\star \geqslant r_{k+1}$, while $h(k)$ is the highest index job $c \leq j$ such that for all $k < \ell \leq c\}$, $d_c - (w_\ell + \ldots + w_c)/s^\star \leqslant d_{\ell-1}$.*

Lemma 5. *Any suffix (a, b) satisfies $b = f(a)$ and any prefix (k, c) satisfies $c = h(k)$.*

The function f requires a little more attention. Since by Lemma 3, the job $a-1$ cannot be scheduled with higher than critical speed, we can assume that a suffix (a, b) is such that a is the smallest index job with $f(a) = b$. So from now on we restrict the domain of f to those jobs. This allows f to be invertible, i.e. $a = f^{-1}(b)$. Note that by definition of f, the job j is in the co-domain of f, meaning that $f^{-1}(j)$ is defined.

Fig. 2. Illustration of a suffix (a, b) and a prefix $(b+1, c)$ in a schedule. Note that job a starts at its release time and job c ends at its deadline.

6. The dynamic program

For every subinstance (i, j), we denote by $Y_{i,j}$ the minimum c_{speed} cost plus $g(r_{j+1} - d_{i-1})$, and by $O_{i,j}$ the minimum $c_{\text{speed}} + c_{\text{mode}}$ cost. If subinstance (i, j) is infeasible we set $Y_{i,j}$, $O_{i,j}$ to $+\infty$. For convenience we denote $g^\star := (g + (s^\star)^\alpha)/s^\star$.

Theorem 1. *The value $O_{i,j}$ satisfies the following recursion. If $j = i - 1$, then $O_{i,j} = \min\{L, g(r_{j+1} - d_{i-1})\}$, otherwise, let $k = f^{-1}(j)$.*

$$O_{i,j} = \min \begin{cases} Y_{i,j} \\ L + g^\star(w_i + \ldots + w_{h(i)}) + O_{h(i)+1,j} \\ Y_{i,k-1} + g^\star(w_k + \ldots + w_j) + L \\ \min Y_{i,a-1} + g^\star(w_a + \ldots + w_b) + L + g^\star(w_{b+1} + \ldots + w_c) + O_{c+1,j}, \end{cases} \quad (1)$$

where the inner minimization is over all jobs $a \in \{i+1, \ldots j\} \cup \{b, c\}$ with $b = f(a)$, $b < j$ and $c = h(b+1)$. As usual if there are no such jobs, the value of this inner minimization is $+\infty$.

Proof. The case $j = i - 1$ is simple, since the optimal empty schedule is either idle or shutdown depending on the span of $[r_{j+1}, d_{i-1})$.

Now for some $i \leq j$, consider the subinstance (i, j). By induction on $j - i$, we can show that for each of the four cases in Eq. (1) there is a feasible schedule with the corresponding cost. For the remainder of the proof, we consider a schedule S minimizing $c_{\text{speed}} + c_{\text{mode}}$ for this subinstance, and we show that one of the four cases yields its cost.

If S is never power down, then the contribution of c_{mode} is exactly $g(r_{j+1} - d_{i-1})$, and the contribution of c_{speed} is minimal. So the first case applies.

Now suppose that there is some interval $[t, u)$ where the schedule is power down, $[t, u)$ is inclusion-wise maximal and it is the first interval. There are several cases now, depending on the conditions $t = \min I$, $u = \max I$, where I is the interval associated to the sub-instance (i, j).

It cannot be that both conditions are true, since this means that the schedule is empty, which contradicts the case assumption $i \leq j$.

If $t = \min I$ and $u < \max I$, then by Lemma 4 there is a prefix (i, c) of the form $c = h(i)$. The portion up to d_c of this schedule has a contribution to cost equal to $L + g^\star(w_i + \ldots + w_c)$, and by the composition of schedules, its remainder has a contribution of $O_{c+1,j}$. Hence, the second case of Eq. (1) applies.

If $t > \min I$ and $u = \max I$, similarly there is a suffix (k, j) and the cost of the schedule up to r_k is $Y_{i,k-1}$, since there are no power down states, while the remainder contributes a cost of $g^\star(w_k + \ldots + w_j) + L$. This time it is the third case of Eq. (1) which applies.

If $t > \min I$ and $u < \max I$, again by Lemma 4, there is a suffix (a, b) and a prefix $(b+1, c)$ around a power down interval $[t, u)$, and by Lemma 5 we have $b = f(a)$, $c = h(b+1)$. Then, the cost of the schedule decomposes into the cost $Y_{i,a-1}$ for the part before r_a, since it does not contain power down states, the cost $g^\star(w_a + \ldots + w_c) + L$ for the part in $[r_a, d_c)$, and the cost $O_{c+1,j}$ for the remainder, by the composition of schedules. In this final case, the last case of Eq. (1) applies. □

7. Complexity analysis

The dynamic program uses $O(n^2)$ variables, and for each one of them a minimization over $O(n)$ values is required. Therefore, it can be run in $O(n^3)$ time.

For a fixed subinstance (i, j) the functions f, h can be computed by simple scanning procedures in linear time as following (we omit their proof of correctness).

- Initially $\ell := i$ and $t := r_i$. For all $k = i, i+1, \ldots, j$, if $t < r_k$, then $\ell := k$, $t := r_k$. In any case $f(\ell) := k$, $t := t + w_k/s^\star$.
- Initially $\ell := j$ and $t := d_j$. For all $k = j, j-1, \ldots, i$, if $t > d_k$, then $\ell := k$, $t := d_k$. In any case $h(k) = \ell$, $t := t - w_k/s^\star$.

The computation of the values $Y_{i,j}$ however is crucial, there are $O(n^2)$ of them and the best known algorithm to compute the optimal schedule for each of them runs in time $O(n^2)$ [6], which would lead to a total running time of $O(n^4)$. We now describe a procedure which permits to compute $Y_{i,j}$ iteratively from $Y_{i-1,j}$ in total time $O(n^2)$. Therefore we can compute all optimal c_{speed} subschedules in total time $O(n^3)$.

7.1. Computing $Y_{i,j}$

The general outline is as follows. We first compute $Y_{1,n}$ in time $O(n^2)$ using the algorithm from Ref. [6]. Then in a first right to left scan we compute all values $Y_{1,j}$ for $j = n-1, \ldots, 1$. After that for every j, we apply a left to right scan to compute all values $Y_{i,j}$ for $i = 2, \ldots, j$. This left to right scan works as follows.

It receives as input the c_{speed}-optimal schedule S for the subinstance $(1, j)$, and applies the following *squeezing* procedure to S. The schedule S consists of a sequence of *blocks*, every block spans some time interval $[t, u)$ and contains a sequence of jobs running at some constant, but block dependent speed.

During the procedure we keep track of the first block which spans time interval $[t, u)$ and schedules the jobs $i, \ldots, b$ at speed s. Initially $i = 1$. We consider the action of squeezing the block to the interval $[u - \ell, u)$ by increasing the speed s, where $\ell := u - (w_i + \ldots + w_b)/s$.

While $i \leq j$, we decide which of the following events happens first, and execute the corresponding actions. See Figure 3.

unfeasibility event: It happens when $d_{i-1} = d_i$ or $r_j = r_{j-1}$. Since in the subinstance (i, j) all jobs are restricted to the interval $[d_{i-1}, r_{j+1}]$, it follows that one of the jobs i, j is restricted to an empty interval, and cannot be scheduled with finite speed. In this case, we announce that subinstance (i, j) is unfeasible, we remove job i from S, and increase i.

merge event: It happens when the current speed s equals speed(u). In this case we merge the first two blocks. (Note that if $u = d_b$, then this event will immediately be followed by the next split event for the merged block.)

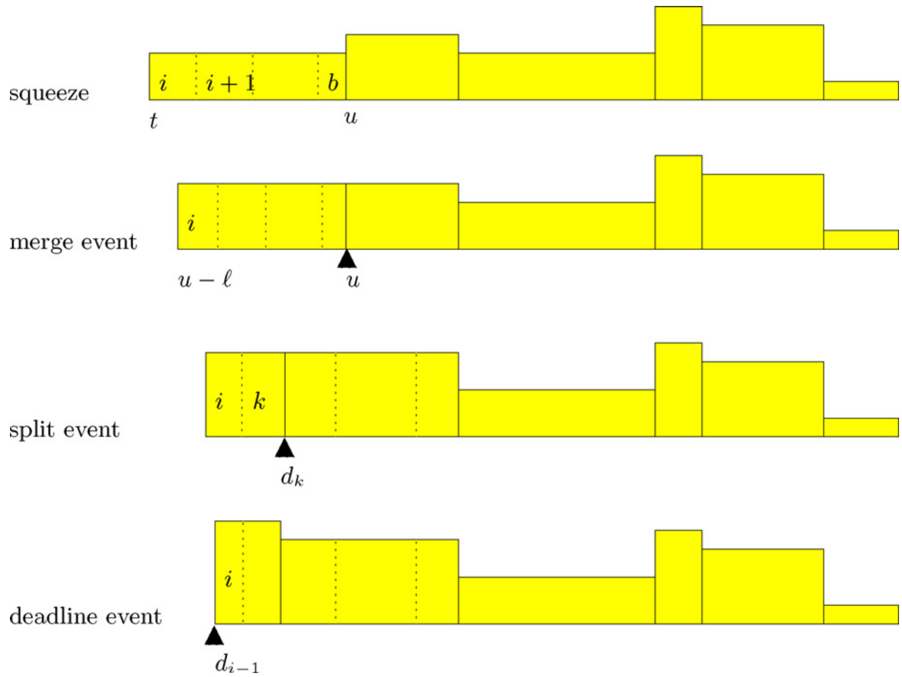

Fig. 3. Different events during the squeeze procedure.

split event: At some moment, a job $i \le k < b$ from the first block might complete at its deadline. This happens when the speed s reaches $\hat{s}(k, b, u) := (w_{k+1} + \ldots + w_b)/(u - d_k)$. In this case the block splits into two new blocks with the first of them restricted to the interval $[t, d_k)$ and to the jobs $i, \ldots, k$.
deadline event: When $s = (w_i + \ldots + w_b)/(u - d_{i-1})$, the current schedule S is the optimal c_{speed}-schedule for the subinstance (i, j). In this case we output S as $Y_{i,j}$, we remove job i from S, and increase i.

At any moment the algorithm maintains a schedule S for the subinstance consisting of all jobs $i, \ldots, j$ with release times and deadlines restricted to the interval $[u - \ell, r_{j+1}]$. We omit the proof of optimality of S which should be straightforward.

It remains to specify how the next event can be determined in constant time. The merge and deadline events, are both specified by a single expression determining the value ℓ at which they occur. For the split event the situation is more subtle, since there are $b - i$ candidates $\hat{s}(k, b, u)$, one for each job $i \le k < b$. We handle this by precomputing $\hat{s}$. Note that for a given job b, there are only $O(n)$ different times u to be considered, and they are of the form d_b, r_{b+1} and r_{j+1} for all $1 \le j \le n$. This is because every block of an optimal schedule ends either at the end of the interval I if it is the last block, or at one of d_b, r_{b+1}, depending on whether the next block has lower or higher speed.

This means that there are $O(n^3)$ values of the form $\hat{s}(k, b, u)$ to compute, and this can be done for each pair b, u in linear time, by iterating k from $b - 1$ to 1. In the procedure above we need to determine the job k, $i \le k < b$ minimizing $\hat{s}(k, b, u)$. Clearly, such a job k can be computed in constant time for each triplet (i, b, u), again by iterating i from $b - 1$ to 1 for each pair b, u.

In the event loop described above every job is responsible for at most three events. Therefore its complexity is $O(n)$ for fixed j, which yields to a total running time of $O(n^3)$.

8. Conclusion

We provided a polynomial time algorithm for the speed scaling with power down scheduling problem, for the special case of agreeable deadlines. This assumption leads to strong structural properties of optimal schedules, which are non-preempted and, moreover, permit a partitioning leading to a dynamic programming algorithm. So the proposed algorithm could not be generalized to instances with arbitrary deadlines. However, we believe that the squeezing procedure could be of independent interest.

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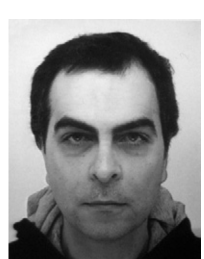

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