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A note on on-line broadcast scheduling with deadlines

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\textbf{A B S T R A C T}

In this paper, we study an on-line broadcast scheduling problem with deadlines, in which the requests asking for the same page can be satisfied simultaneously by broadcasting this page, and every request is associated with a release time, deadline and a required page with a unit size. The objective is to maximize the number of requests satisfied by the schedule. In this paper, we focus on an important special case where all the requests have their spans (the difference between release time and deadline) less than 2. We give an optimal online algorithm, i.e., its competitive ratio matches the lower bound of the problem.

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\section{1. Introduction}

Broadcasting technologies receive a lot attention on networks that employs broadcasting to disseminate data or information. In contrast to the traditional point-to-point mode of communication, broadcasting technologies have an advantage that one broadcast by the server can simultaneously satisfy requests required from multiple clients for an identical message. In this paper, we focus a \textit{pull-based} model of broadcast scheduling problems, which is formalized as below.

\textbf{Problem description.} There is a collection of pages \(S = \{1, \ldots, n\}\), in the server. The clients send requests to ask for these pages and each request has a release time, deadline and a distinct page to ask for. The server answers requests by broadcasting pages. Note that a broadcast of a page can satisfy all the requests asking for the same page simultaneously and there is at most one page to be broadcasted at any time. During broadcasting, the preemption is allowed, but if the broadcast of a page is preempted, then in case the server choose this page to broadcast again, it must from the start point not the break point, we call this as \textit{preemption with restart}. When the request for the page that currently broadcast arrives it must be kept in the queue of unsatisfied requests.

In this paper, we consider each page has an unit size, i.e., any page can be broadcasted during one time unit, and the requests arrive over time. The scheduling algorithm used by the server has no knowledge of requests in advance and makes decisions only with information of requests having already arrived. We call this on-line broadcast scheduling. There are two models, discrete and continuous models. In discrete model, the arrival times and deadlines for all the requests are integral. For the online version of discrete model, Kim and Chwa \cite{kim} gave a best possible online algorithm with competitive ratio 2. Since new requests may arrive and end any time, the continuous model is more general, and is well-studied during these years. The main objectives are minimizing the flow time (response time) and maximizing the total throughput, i.e., the number of satisfied requests.

\textbf{Previous results.} Most of the previous works on on-line broadcast scheduling focus on minimizing the flow time \([1,4,6-8,11]\). On maximizing the throughput of on-line broadcast scheduling, Kim and Chwa \cite{kim} first gave a 5.828-competitive algorithm, then Chan et al. \cite{chan} showed that the competitive ratio of algorithm in \cite{kim} is at most 5. Recently, Zheng et al. \cite{zheng} obtained a new on-line algo-
algorithm by looking forward two steps and proved that the
competitive ratio is at most 4.56. The lower bound of
maximizing the throughput of on-line broadcast scheduling
is 4 which is from a related on-line interval scheduling
problem [12]. Fung et al. [9] first studied this on-line
broadcast scheduling problem with laxity (the span of a
request, i.e., the difference between deadline and release
time) constraints, and got some results as below. If all the
requests have their laxity at least 2 then a nice and simple
online algorithm with competitive ratio 2.618 is given. If
all the requests have their laxity at most $\alpha < 2$, then an
$f(\alpha)$-competitive algorithm can be achieved, where
$4 < f(\alpha) \leq 4.714$. For off-line broadcast scheduling problem,
maximizing the throughput is NP-hard [4].

Our contributions. In this paper, we focus on maximizing
the throughput and give a 4-competitive algorithm if all the
requests have laxity less than 2, which is optimal since the
lower bound of this problem is also 4.

2. Preliminaries

A request $R_i$ is defined as a triple $(p_i, r_i, d_i)$, where $p_i$
is the requested page, $r_i$ and $d_i$ are its release time and
deadline, respectively.

Definition 1 (Laxity). For a request $R = (p, r, d)$, its laxity
is defined as $(d - r)$.

Definition 2 (Alive and dead). Given a request $R = (p, r, d)$,
if $(d - t) \geq 1$ then we say the request is alive at time $t$, otherwise, dead at time $t$.

For a request $R = (p, r, d)$, at time $t$ its weight $W(R, t)$
is defined as the following table, i.e., if it is alive at time $t$
then its weight is 1 otherwise 0.

<table>
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<tr>
<th>$d - t$</th>
<th>$(-\infty, 1]$</th>
<th>$[1, +\infty)$</th>
</tr>
</thead>
<tbody>
<tr>
<td>Weight</td>
<td>0</td>
<td>1</td>
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For a page $P$, its weight $W(P, t)$ is defined as the number
of all requests alive at time $t$ in a pending list, i.e.,

$$W(P, t) = \sum_{p_i = P} W(R_i, t).$$

Definition 3 (Competitive ratio). To evaluate an online algo-
rithm, we use the standard measure called competitive
ratio. For any input sequence $L$, let $A(L)$ be the cost by an
online algorithm $A$ and $OPT(L)$ be the cost by an optimal
off-line algorithm. The competitive ratio of algorithm $A$ is
then defined as $\rho_A = \sup_L \frac{A(L)}{OPT(L)}$.

3. A tight upper bound for laxity less than 2

We first give an on-line algorithm then show that its
competitive ratio is 4 which matches the lower bound [12].
Our algorithm is quite similar with ones in [5,10]. The
main ideas of our algorithm are: (i) whenever we decide
to broadcast a page, the page with the maximal weight is
selected to be served; (ii) when a new request for page $p_0$
arrives if to start broadcasting page $p_0$ can double the
profit (i.e., throughput), then we abort broadcasting the
current page, and put the new request into a pending list
and select a page with the maximal weight and broadcast
that page (this is the difference between our algorithm and
the ones in [5,10]). Otherwise, continue to broadcast the
current page and put the new request in the pending list.

Let $P_0$ be the page of a new request which arrives at
the current time $t$, let $P_1$ denote the currently broadcasted
page if it exists. Our algorithm is described (see Algo-
rithm 1).

We first define a concept called basic chain and observe
an important property related to it. Then, we divide the
broadcasts by our algorithm into a set of basic chains and
combine the property to get an upper bound 4 for the
competitive ratio.

Definition 4 (Basic chain). For $i \leq j$, a sequence of broad-
casts $(P_i, P_{i+1}, \ldots, P_j)$ is called a basic chain if pages
$P_i, \ldots, P_{j-1}$ are aborted broadcasts and page $P_j$ is a
completed broadcast, and the broadcast just before $P_j$ is empty
or a completed broadcast.

Theorem 1. For any input list of requests with laxity less than 2,
the competitive ratio of our algorithm is 4.

Proof. Let the sequence of broadcasts $(P_1, P_2, \ldots, P_j)$ be
the first basic chain generated by our on-line algorithm.
Let $(P_1^*, P_2^*, \ldots, P_m^*)$ be the first pages broadcast by an
optimal scheduling such that the starting point of broad-
casting page $P_m^*$ is sat in the time interval for broad-
casting page $P_j$, shown as Fig. 1 (if $P_m^*$ does not exist,
then we set $P_m^*$ as a dummy page). Let time $t_i^*$ ($t_j^*$)
denote the starting point of broadcasting page $P_i$ ($P_j^*$) for
$1 \leq i \leq m$ and $t_i^*$ denotes the starting point of broadcasting page $P_i$ ($P_j^*$) if
$1 \leq i \leq m$. Without loss of generality, assume that $t_i^* - t_j^* > 1$ for
$1 \leq i \leq m$ otherwise we can get another
optimal schedule by broadcasting $P_i^*$ at time $t_i^*$ except for
page $P_j$ and doing nothing during $[t_j^*, t_j^*+1]$. Now,
for $1 \leq i \leq j$, we define a set of intervals,
$$I_i = [t_i, t_i+1], \quad \text{where } t_i+1 = t_j + 1.$$ For an input list $L$,
let $A(L)$ and $OPT(L)$ be the number of requests satisfied
by our algorithm and an optimal schedule respectively. For
$1 \leq i \leq m$, let $x_i^*$ be the number of requests which are satis-
ified by broadcasting page $P_i^*$ in the optimal schedule.
Remember that the whole sequence of broadcasts \((P_1, P_2, \ldots, P_j)\) by our algorithm is a basic chain. Then we have the following observation.

**Claim 1.** For \(1 \leq i \leq m\), \(x_i^m \leq 2^{j+1-m}W\), where \(W = W(P_j, t_j)\).

**Proof.** Recall that \(t_i^*\) is the start time to broadcast page \(P_i^*\). For \(1 \leq i \leq m\), we have

\[
(t_{i+1}^* - t_i^*) \geq 1. \quad (1)
\]

On the other hand, the whole sequence of broadcasts \((P_1, P_2, \ldots, P_j)\) by our algorithm is a basic chain, by the definition of a basic chain, for \(1 \leq i \leq j\) we have

\[
t_{i+1} - t_i \leq 1. \quad (2)
\]

Let \(I_f = [t_f, tf+1)\) be the interval such that \(t_i^* \in I_f\). By (1) and (2), we have the number of broadcasts during \([t_f, tf]\) by our algorithm is not less than the number by the optimal schedule during \([t_i^*, t_m]\), i.e.,

\[
(j - f) \geq (m - i) \implies f \leq j - (m - i) = j + i - m. \quad (3)
\]

According to algorithm WP, for any page \(P\), we have

\[
W(P, t_f) \leq 2^{f-j}W(P_j, t_j) = 2^{f-j}W. \quad (4)
\]

Let \(t'_i\) be the time when the last request for page \(P_i^*\) released at or before time \(t_i^*\) in the input list. Let \(t'_i = \max\{t_f, t_i^*\}\). According to algorithm WP, at time \(t'_i\) there is a comparison between \(W(P_i^*, t_i^*)\) and \(W(P_j, t_f)\). Since page \(P_j\) is broadcast at time \(t_j^*\) by our algorithm, we have

\[
W(P_i^*, t_i^*) < 2W(P_j, t_f). \quad (5)
\]

By inequalities (3)-(5), we have

\[
W(P_i^*, t_i^*) \leq 2^{j+1-m}W. \quad (6)
\]

Let \(x_i^m\) be the number of requests for page \(P_i^*\) which are alive at time \(t_i^m\). By the definition of \(x_i^m\) and \(t_i^m\), we have

\[
x_i^m \leq x_i^m. \quad (7)
\]

By (7) and (6)

\[
x_i^m \leq W(P_i^*, t_i^m) \leq 2^{j+1-m}W. \quad (8)
\]

Next we are going to bound \(\text{OPT}(L)\), i.e., the throughput by the optimal schedule. Without loss of generality, broadcasts generated by our algorithm and the optimal algorithm looks like Fig. 1.

We prove this theorem by mathematical induction over the number of basic chains produced by our algorithm.

First, we prove that the theorem holds if the whole broadcast generated by algorithm WP is a basic chain. Then we assume that the theorem holds for the case in which there are \(h\) basic chains. Finally, we prove that the theorem still holds for \((h+1)\) basic chains.

**Step 1.** There is only one basic chain in the whole broadcast generated by our algorithm. In this case, we prove \(P_m^*\) is the last page in the optimal schedule. Otherwise there is at least one request alive at time \(t_m^* + 1\), where \(t_m^* \geq t_j\). Since all the requests have laxity less than 2, the alive request must be released after \(t_j\). So, the request would have been broadcast by our algorithm after \(t_j + 1\). But, this contradicts with the fact that the whole sequence of broadcasts \((P_1, P_2, \ldots, P_j)\) generated by our algorithm is one basic chain. So page \(P_m^*\) is the last page in the optimal schedule. By Claim 1, we have

\[
\text{OPT}(L) = \sum_{i=1}^{m} x_i^m \leq \sum_{i=1}^{m} 2^{j+1-m}W
\]

\[
= 4W \sum_{i=0}^{m-1} 2^{i} \leq 4W = 4A(L).
\]

**Step 2.** Assume that this theorem holds when there are \(h\) basic chains in the whole broadcast generated by our algorithm, where \(h \geq 1\). Next we consider the case in which there are \(h+1\) basic chains in our broadcast. First, we define four sublists of requests. We define \(L_2\) as the sublist of requests that will be considered by our algorithm after time \(t_j + 1\), i.e., the sublist of requests with release time at least \(t_j + 1\) or requests which are still alive at time \(t_j + 1\) \((t_m^* + 1)\) and not satisfied by our algorithm before \(t_j + 1\). In the same way, we define \(L_2^*\) as the sublist of requests that will be considered by the optimal schedule after \(t_m^* + 1\), i.e., \(L_2^*\) is the sublist of requests with release time at least \(t_m^* + 1\) or requests which are still alive at time \(t_m^* + 1\) and not satisfied by the optimal algorithm before \(t_m^* + 1\). Let \(L_1 = L - L_2\) and \(L_2^* = L - L_2^*\). By definitions, \(L_1\) is the sublist of requests with release time before \(t_j + 1\) and not alive at time \(t_j + 1\), i.e., requests with release time before \(t_j + 1\) and not satisfied, or requests satisfied before \(t_j + 1\). Observe that for any request \(R \in L_1\), if request \(R\) is not satisfied by algorithm WP, then request \(R\) is not alive at time \(t_j + 1\), therefore \(R\) is not alive at time \(t_m^* + 1\) too, where \(t_m^* \geq t_j\). Then \(R \notin L_2^*\). If request \(R\) is satisfied by algorithm WP before \(t_j + 1\), then it is not alive at time \(t_m^* + 1\) since every request has laxity less than 2. So, in both cases, we have \(R \notin L_2^*\), where \(R \in L_1\). Then

\[
L_2^* \subseteq L_2 = L - L_1. \quad (9)
\]

To estimate the throughput by algorithm WP and the optimal algorithm, we need to modify \(L_2\) and \(L_2^*\) slightly. For every request in \(L_2^*\) \((\in L_2^*)\), if its release time is at least
t_j + 1 then just copies it into L_1^2 (L_2^{*1}) else modifies its release time to t_j + 1 then copies it into L_2^2 (L_2^{*1}). For every request in L_2^*, if its release time is at least t_m + 1 then just copies it into L_2^2 else modifies its release time to t_m + 1 then copies it into L_2^2. By the above definitions and (9), we have

\[ \text{OPT}(L_1^2) \geq \text{OPT}(L_2^{*1}) \geq \text{OPT}(L_2^2). \]  \hfill (10)
\[ A(L) = A(L_1) + A(L_2^1). \]  \hfill (11)
\[ \text{OPT}(L) = \text{OPT}(L_1^*) + \text{OPT}(L_2^2). \]  \hfill (12)

And by the assumption for h basic chains and Claim 1, we have

\[ 4A(L_1^1) \geq OPT(L_2^1) \quad \text{and} \quad 4A(L_1) \geq OPT(L_1^*). \]  \hfill (13)

So,

\[ 4A(L) = 4(A(L_1) + A(L_2^1)) \quad \text{by (11)} \]
\[ \geq \text{OPT}(L_1^*) + \text{OPT}(L_2^1) \quad \text{by (13)} \]
\[ \geq \text{OPT}(L_1^*) + \text{OPT}(L_2^*2) \quad \text{by (10)} \]
\[ = \text{OPT}(L) \quad \text{by (12)} \]

Hence, this theorem holds. □

References