Precedence-constrained scheduling problems parameterized by partial order width

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Abstract. Negatively answering a question posed by Mnich and Wiese (Math. Program. 154(1-2):533-562), we show that $P2|\text{prec}, p_j \in \{1,2\}|C_{\text{max}}$, the problem of finding a non-preemptive minimum-makespan schedule for precedence-constrained jobs of lengths 1 and 2 on two parallel identical machines, is W[2]-hard parameterized by the width of the partial order giving the precedence constraints. To this end, we show that Shuffle Product, the problem of deciding whether a given word can be obtained by interleaving the letters of k other given words, is W[2]-hard parameterized by k, thus additionally answering a question posed by Rizzi and Vialette (CSR 2013). Finally, refining a geometric algorithm due to Servakh (Diskretn. Anal. Issled. Oper. 7(1):75–82), we show that the more general Resource-Constraint Project Scheduling problem is fixed-parameter tractable parameterized by the partial order width combined with the maximum allowed difference between the earliest possible and factual starting time of a job.

Keywords: resource-constrained project scheduling, parallel identical machines, makespan minimization, parameterized complexity, shuffle product

1 Introduction

We study the parameterized complexity of the following NP-hard problem and various special cases [13, 15] with respect to the width of the given partial order.

Problem 1.1 (Resource-constrained project scheduling (RCPSP)).

Input: A set J of jobs, a partial order \leq on J, a set R of renewable resources, for each resource $\rho \in R$ the available amount R_{ρ} , and for each $j \in J$ a processing time $p_j \in \mathbb{N}$ and the amount $r_{j\rho} \leq R_{\rho}$ of resource $\rho \in R$ that it consumes.

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Find: A schedule $(s_i)_{i \in J}$, that is, a starting time $s_i \in \mathbb{N}$ of each job j, such that

- 1. for $i \prec j$, job i finishes before job j starts, that is, $s_i + p_i \leq s_j$,
- 2. at any time t, at most R_{ρ} units of each resource ρ are used, that is, $\sum_{j \in s(t)} r_{j\rho} \leq R_{\rho}$, where $s(t) := \{j \in J \mid t \in [s_j, s_j + p_j)\}$, and
- 3. the maximum completion time $C_{\max} := \max_{j \in J} (s_j + p_j)$ is minimum. A schedule satisfying (1)–(2) is *feasible*; a schedule satisfying (1)–(3) is *optimal*.

Intuitively, a schedule $(s_j)_{j\in J}$ processes each job $j\in J$ non-preemptively in the half-open real-valued interval $[s_j,s_j+p_j)$, which costs $r_{j\rho}$ units of resource ρ during that time. After finishing, jobs free their resources for later jobs. If there is only one resource and each job j requires one unit of it, then RCPSP is equivalent to P|prec| C_{max} , the NP-hard problem of non-preemptively scheduling precedence-constrained jobs on a given number m of parallel identical machines to minimize the maximum completion time [15].

Mnich and Wiese [11] asked whether $P|\operatorname{prec}|C_{\max}$ is solvable in $f(p_{\max}, w)$ poly(n) time, where p_{\max} is the maximum processing time, w is the width of the given partial order \leq , n is the input size, and f is a computable function independent of the input size. In other words, the question is whether $P|\operatorname{prec}|C_{\max}$ is fixed-parameter tractable parameterized by p_{\max} and w. Motivated by this question, which we answer negatively, we strengthen hardness results for $P|\operatorname{prec}|C_{\max}$ and refine algorithms for RCPSP with small partial order width.

Due to space constraints, some details are deferred to an Appendix.

Stronger hardness results. We obtain new hardness results for the following special cases of P|prec| C_{max} (for basic definitions of parameterized complexity terminology, see the end of this section and recent textbooks [4, 5]):

- (1) P2|chains| C_{max} , the case with two machines and precedence constraints given by a disjoint union of total orders, remains weakly NP-hard for width 3.
- (2) P2|prec, $p_j \in \{1, 2\} | C_{\text{max}}$, the case with two machines and processing times 1 and 2, is W[2]-hard parameterized by the partial order width w.
- (3) P3|prec, $p_j=1$,size $_j \in \{1,2\}|C_{\max}$, the case with three machines, unit processing times, but where each job may require one or two machines, is also W[2]-hard parameterized by the partial order width w.

Towards showing (2) and (3), we show that Shuffle Product, the problem of deciding whether a given word can be obtained by interleaving the letters of k other given words, is W[2]-hard parameterized by k. This answers a question of Rizzi and Vialette [12]. We put these results into context in the following.

Result (1) complements the fact that $P|\operatorname{prec}|C_{\max}$ with constant width w is solvable in pseudo-polynomial time using dynamic programming [14] and that $P2|\operatorname{chains}|C_{\max}$ is $\operatorname{strongly}$ NP-hard for unbounded width [6].

Result (2) complements the NP-hardness result for P2|prec, $p_j \in \{1,2\}|C_{\max}$ due to Ullman [15] and the W[2]-hardness result for P|prec, $p_j = 1|C_{\max}$ parameterized by the number m machines due to Bodlaender and Fellows [3]. While not made explicit, one can observe that Bodlaender and Fellows' reduction creates hard instances with w = m + 1. This is remarkable since P|prec| C_{\max} is trivially polynomial-time solvable if $w \leq m$, and also since the result negatively

answered Mnich and Wiese's question [11] twenty years before it was posed. Our result (2), however, gives a stronger negative answer: unless W[2] = FPT, not even P2|prec, $p_i \in \{1,2\}|C_{\text{max}}$ allows for the desired $f(w) \cdot \text{poly}(n)$ -time algorithm.

Refined algorithms. Servakh [14] gave a geometric pseudo-polynomial-time algorithm for RCPSP with constant partial order width w. The degree of the polynomial depends on w and, by (1) above, the algorithm cannot be turned into a true polynomial-time algorithm unless P = NP even for constant w. We refine this algorithm to solve RCPSP in $(2\lambda+1)^w \cdot 2^w \cdot \text{poly}(n)$ time, where λ is the maximum allowed difference between earliest possible and factual starting time of a job. The degree of the polynomial depends neither on w nor λ and is indeed a polynomial of the input size n. This does not contradict (1) since the factor $(2\lambda+1)^w$ might be superpolynomial in n. We note that fixed-parameter tractability for w or λ alone is ruled out by (2) and by Lenstra and Rinnooy Kan [10], respectively.

Preliminaries. A reflexive, symmetric, and transitive relation \preceq on a set X is a partial order. We write $x \prec y$ if $x \preceq y$ and $x \neq y$. A subset $X' \subseteq X$ is a chain if \preceq is a total order on X'; it is an antichain if the elements of X' are mutually incomparable by \preceq . The width of \preceq is the size of largest antichain in X. A chain decomposition of X is a partition $X = X_1 \uplus \cdots \uplus X_k$ such that each X_i is a chain.

Recently, the parameterized complexity of scheduling problems attracted increased interest [2]. The idea is to accept exponential running times for solving NP-hard problems, but to restrict them to a small parameter [4, 5]. Instances (x, k) of a parameterized problem $\Pi \subseteq \Sigma^* \times \mathbb{N}$ consist of an input x and a parameter k. A parameterized problem Π is fixed-parameter tractable if it is solvable in f(k) · poly(|x|) time for some computable function f. Note that the degree of the polynomial must not depend on k. FPT is the class of fixed-parameter tractable parameterized problems. There is a hierarchy of parameterized complexity classes FPT \subseteq W[1] \subseteq W[2] $\subseteq \cdots \subseteq$ W[P], where all inclusions are conjectured to be strict. A parameterized problem Π_2 is W[t]-hard if there is a parameterized reduction from each problem $\Pi_1 \in$ W[t] to Π_2 , that is, an algorithm that maps an instance (x, k) of Π_1 to an instance (x', k') of Π_2 in time f(k) · poly(|x|) such that $k' \leq g(k)$ and $(x, k) \in \Pi_1 \Leftrightarrow (x', k') \in \Pi_2$, where f and g are arbitrary computable functions. No W[t]-hard problem is fixed-parameter tractable unless FPT = W[t].

2 Parallel identical machines and shuffle products

This section presents our hardness results for special cases of P|prec| C_{max} . In Section 2.1, we show weak NP-hardness of P2|chains| C_{max} for three chains. In Section 2.2, we show W[2]-hardness of Shuffle Product as a stepping stone towards showing W[2]-hardness of P3|prec, p_j =1,size $_j$ \in {1, 2}| C_{max} and P2|prec, p_j \in {1, 2}| C_{max} parameterized by the partial order width in Section 2.3.

2.1 Weak NP-hardness for two machines and three chains

Du et al. [6] showed that P2|chains| C_{max} is strongly NP-hard. We complement this result by the following theorem.

$s_1 =$	a			c		b				b
	- 1			1		1				1
$s_2 =$	- 1		b	- 1	b	1	c			- 1
	- 1		- 1	- 1	1	1	1			- 1
$s_3 =$	- 1	c	- 1	- 1	1	- 1	1	a	b	- 1
- 0	Ţ	₩	*	Ţ	*	4	↓	4	~	4
t =	a	c	b	c	b	b	c	a	b	b

Fig. 2.1. Illustration of a shuffle product: for $s_1 = acbb$, $s_2 = bbc$, and $s_3 = cab$, one has $t = acbcbbcabb \in s_1 \sqcup s_2 \sqcup s_3$. Dashed arcs show how the letters of each s_i map into t.

Theorem 2.1. P2|chains| C_{max} is weakly NP-hard even for precedence constraints of width three, that is, consisting of three chains.

Proof (sketch). We reduce from the weakly NP-hard Partition problem [8, SP12]: Given a multiset of positive integers $A = \{a_1, \ldots, a_t\}$, decide whether there is a subset $A' \subseteq A$ such that $\sum_{a_i \in A'} a_i = \sum_{a_i \in A \setminus A'} a_i$. Let $A = \{a_1, \ldots, a_t\}$ be a Partition instance. If $b := \left(\sum_{a_i \in A} a_i\right)/2$ is not an integer, then we are facing a no-instance. Otherwise, we construct a P2|chains| C_{\max} instance as follows. Create three chains $J^0 := \{j_1^0 \prec \cdots \prec j_t^0\}$, $J^1 := \{j_1^1 \prec \cdots \prec j_{t+1}^1\}$, and $J^2 := \{j_1^2 \prec \cdots \prec j_{t+1}^2\}$ of jobs. For each $i \in \{1, \ldots, t\}$, job j_i^0 gets processing time a_i . The jobs in $J^1 \cup J^2$ get processing time 2b each. This construction can be performed in polynomial time and one can show that the input Partition instance is a yes-instance of and only if the created P2|chains| C_{\max} instance allows for a schedule with makespan T := (2t+3)b: in such a schedule, each machine must perform exactly t+1 jobs from $J^1 \cup J^2$ and has b time for jobs from J^0 . \square

2.2 W[2]-hardness for Shuffle Product

In this section, we show a W[2]-hardness result for Shuffle Product that we transfer to P2|prec, $p_j \in \{1,2\} | C_{\text{max}}$ and P3|prec, $p_j = 1$,size $_j \in \{1,2\} | C_{\text{max}}$ in Section 2.3. We first formally introduce the problem (cf. Figure 2.1).

Definition 2.2 (shuffle product). By s[i], we denote the ith letter in a word s. A word t is said to be in the *shuffle product* of words s_1 and s_2 , denoted by $t \in s_1 \sqcup s_2$, if t can be obtained by interleaving the letters of s_1 and s_2 . Formally, $t \in s_1 \sqcup s_2$ if there are increasing functions $f_1: \{1, \ldots, |s_1|\} \to \{1, \ldots, |t|\}$ and $f_2: \{1, \ldots, |s_2|\} \to \{1, \ldots, |t|\}$ mapping positions of s_1 and s_2 to positions of t such that, for all $i \in \{1, \ldots, |s_1|\}$ and $j \in \{1, \ldots, |s_2|\}$, one has $t[f_1(i)] = s_1[i]$, $t[f_2(j)] = s_2[j]$, and $f_1(i) \neq f_2(j)$. This product is associative and commutative, which implies that the shuffle product of any set of words is well-defined.

Problem 2.3 ((Binary) Shuffle Product).

Input: Words s_1, \ldots, s_k , and t over a (binary) alphabet Σ .

Parameter: k.

Question: Is $t \in s_1 \sqcup s_2 \sqcup \cdots \sqcup s_k$?

BINARY SHUFFLE PRODUCT is NP-hard for unbounded k [16, Lemma 3.2], whereas SHUFFLE PRODUCT is polynomial-time solvable for constant k using

dynamic programming. Rizzi and Vialette [12] asked about the parameterized complexity of Shuffle Product. We answer the question by the following theorem.

Theorem 2.4. BINARY SHUFFLE PRODUCT is W[2]-hard.

Our proof uses a parameterized reduction from the W[2]-hard DOMINATING SET problem [4, 5] and is inspired by Bodlaender and Fellows's proof that $P|\text{prec}, p_i=1|C_{\text{max}}$ is W[2]-hard parameterized by the number m of machines [3].

Problem 2.5 (Dominating Set).

Input: A graph G = (V, E) and a natural number k.

Parameter: k.

Question: Is there a size-k dominating set D, that is, $V \subseteq N[D]$?

Herein, N[D] is the set of vertices in D and their neighbors. In order to describe the construction, we introduce some notation.

Definition 2.6. We denote the concatenation of words s_1, \ldots, s_k as $\prod_{i=1}^k s_i := s_1 s_2 \ldots s_k$ and denote k repetitions of a word s by s^k . The number of occurrences of a letter a in a word s is $|s|_a$.

Construction 2.7. Given a DOMINATING SET instance (G, k) with a graph G = (V, E), we construct an instance of BINARY SHUFFLE PRODUCT with k+3 words over $\Sigma = \{a, b\}$ in polynomial time as follows. The construction is illustrated in Figure 2.2. Without loss of generality, assume that $V = \{1, \ldots, n\}$.

For
$$u, v \in V$$
, let $\ell_{u,v} := \begin{cases} 1 & \text{if } u = v \text{ or } \{u, v\} \in E, \\ 2 & \text{otherwise.} \end{cases}$ (2.1)

Moreover, define two words

$$A := \prod_{u=1}^{n} \prod_{v=1}^{n} ab^{\ell_{u,v}}$$
 and $B := ((a^k b^{2k})^{n-1} a^k b^{2k-1})^n$.

Finally, let N := 2k(n-1) + 1 and output an instance of Shuffle Product with the following k+3 words:

$$\begin{split} s_i &:= A^N \quad \text{for each } i \in \{1, \dots, k\}, \qquad \qquad t := B^N (a^k b^{2k})^{n-1}, \\ s_{k+1} &:= a^{|t|_a - \sum_{i=1}^k |s_i|_a}, \quad \text{and} \qquad \qquad s_{k+2} := b^{|t|_b - \sum_{i=1}^k |s_i|_b}. \end{split}$$

Note that A is simply the word that one obtains by concatenating the rows of the adjacency matrix of G and replacing ones by ab and zeroes by abb.

Before showing the correctness of Construction 2.7, we make some basic observations about the words it creates, for which we introduce some terminology.

Definition 2.8 (long and short blocks, positions). A block in a word s is a maximal consecutive subword using only one letter. A c-block is a block containing only the letter c. A block has position i in s if it is the ith successive block in s. We call b-blocks of length 2k-1 in t short and b-blocks of length 2k long.

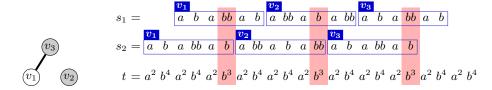


Fig. 2.2. Left: A DOMINATING SET instance with k=2 and a solution $\{v_2, v_3\}$ (the gray nodes). Right: The "base pattern" of the corresponding Shuffle Product instance (only one repetition of A in s_1 and s_2 and only one repetition of B in t is shown). Blocks of s_1 and s_2 are mapped into the blocks of t displayed in the same column. The horizontal (blue) rectangles reflect that each s_i is built as the concatenation of the rows of the adjacency matrix, where zeroes are replaced by abb and ones by ab. The amount of horizontal offset of each s_i corresponds to the selection of a vertex as dominator (v_2 for s_1 and v_3 for s_2). The dark columns (red) correspond to the short b-blocks of t: they ensure that, in each row of the adjacency matrix, at least one selected vertex dominates the vertex corresponding to that row. The base pattern is repeated N times to ensure that at least one occurrence of the pattern is mapped to t without unwanted gaps. Additional words s_{k+1} and s_{k+2} are added to match the remaining letters from t.

Observation 2.9. The words s_1, \ldots, s_k and t created by Construction 2.7 from a DOMINATING SET instance (G, k) have the following properties:

- (i) Each s_i for $i \in \{1, ..., k\}$ contains $2Nn^2$ blocks.
- (ii) The word t contains $2Nn^2 + 2(n-1)$ blocks.
- (iii) For $i \in \{1, ..., k\}$, all a-blocks in s_i have length 1. All a-blocks of t have length k.
- (iv) For $h \in \{1, ..., Nn\}$, the b-blocks at position 2hn in t are short. All other b-blocks in t are long.
- (v) For each $i \in \{1, ..., k\}$, $p \in \{0, ..., N-1\}$, and $u, v \in \{1, ..., n\}$, the b-block at position $2pn^2 + 2n(u-1) + 2v$ in s_i has length $\ell_{u,v}$: it corresponds to the entry in the uth row and vth column of the adjacency matrix of G.

Since Construction 2.7 runs in polynomial time and the number of words in the created Shuffle Product instance only depends on the size of the sought dominating set, for Theorem 2.4, it remains to prove the following lemma.

Lemma 2.10. Let s_1, \ldots, s_{k+2} and t be the words created by Construction 2.7 from a DOMINATING SET instance (G, k). Then G has a dominating set of size k if and only if $t \in s_1 \sqcup s_2 \sqcup \cdots \sqcup s_{k+2}$.

Proof. (\Rightarrow) Assume first that G = (V, E) has a dominating set $D = \{d_1, \ldots, d_k\}$. We describe t as a shuffle product of the words s_i as follows. For each $i \in \{1, \ldots, k\}$, map all letters from the block at position x of s_i into block $x + 2(n - d_i)$ of t, that is, consecutive blocks of s_i are mapped into consecutive blocks of t with a small offset depending on d_i . So far, at most t letters are mapped into each t-block of t and at most t letters are mapped into each t-block of t. Hence, all t-blocks and all long t-blocks of t are long enough to accommodate all their

designated letters. It remains to show that at most 2k-1 letters are mapped into each short b-block β of t. By Observation 2.9(iv), β is at position 2hn for some $h \in \{1, \ldots, Nn\}$. Thus, there are $p \in \{0, \ldots, N-1\}$ and $u \in \{1, \ldots, n\}$ such that $2hn = 2(pn + u)n = 2pn^2 + 2un$. For each s_i , the block α_i of s_i mapped into β has position $(2pn^2 + 2un) - 2(n - d_i) = 2pn^2 + 2(u - 1)n + 2d_i$. Hence, α_i has length ℓ_{u,d_i} by Observation 2.9(v). Since D is a dominating set, it contains a vertex d_{i^*} such that $d_{i^*} = u$ or $\{d_{i^*}, u\} \in E$. Thus, by (2.1), α_{i^*} has length $\ell_{u,d_{i^*}} = 1$. Overall, at most k b-blocks of $\{s_1, \ldots, s_k\}$ are mapped into β . We have shown that at least one of them, namely α_{i^*} , has length one. Since the others have length at most two, at most 2k-1 letters are mapped into block β .

We have seen a mapping of the words s_i with $i \in \{1, \ldots, k\}$ to t. Thus, we have $|t|_a \geq \sum_{i=1}^k |s_i|_a$ and $|t|_b \geq \sum_{i=1}^k |s_i|_b$ and the words s_{k+1} and s_{k+2} are well-defined. It remains to map s_{k+1} and s_{k+2} to t. Since s_{k+1} consists only of a and s_{k+2} only of b, we only have to check that t contains as many letters with letters a or b as all words s_i together, which is true by the definition of s_{k+1} and s_{k+2} . We conclude that $t \in s_1 \coprod s_2 \coprod \cdots \coprod s_{k+2}$ if G has a dominating set of size k.

- (\Leftarrow) Assume that $t \in s_1 \sqcup s_2 \sqcup \cdots \sqcup s_{k+2}$. We show that G has a dominating set of size k. To this end, for $i \in \{1, \ldots, k\}$, let $y_i(x)$ be the position of the block in t into which the last letter of the block at position x of s_i is mapped and let $\delta_i(x) = y_i(x) x$. We will see that, intuitively, one can think of $\delta_i(x)$ as the shift of the xth block of s_i in t. To show that G has a dominating set of size k, we use the following two facts about δ_i , which we will prove afterwards.
 - (i) For $i \in \{1, ..., k\}$ and $x \in \{1, ..., 2Nn^2\}$, one has $\delta_i(x) \in \{0, ..., 2(n-1)\}$.
 - (ii) There is a $p \in \{0, ..., N-1\}$ such that, for all $i \in \{1, ..., k\}$, δ_i is constant over the interval $I_p = \{2pn^2 + 1, ..., 2(p+1)n^2 + 1\}$.

We now focus on a $p \in \{0, ..., N-1\}$ as in (ii) and write δ_i for the value $\delta_i(x)$ taken for all $x \in I_p$. We show that $D := \{d_i = n - \delta_i/2 \mid k \in \{1, ..., k\}\}$ is a dominating set of size k for G, that is, we show $D \subseteq V$ and $V \subseteq N[D]$.

To this end, consider a vertex $u \in V$ and the block β of t at position $2pn^2 + 2un = 2hn$ for $h = pn + u \in \{1, \ldots, Nn\}$. By Observation 2.9(iv), β is a short b-block. For any $i \in \{1, \ldots, k\}$, let α_i be the block at position $2pn^2 + 2un - \delta_i$ in s_i . Because of (i), this position is in I_p . By definition of δ_i , the last letter of α_i is mapped into β . Thus, α_i is a b-block. Note that this implies that δ_i is even since a-blocks and b-blocks are alternating in t and s_i . Moreover, by (i), $d_i = n - \delta_i/2 \in \{1, \ldots, n\} = V$. It follows that $D \subseteq V$. We show that $u \in N[D]$. To this end, note that the a-block in s_i at position $2pn^2 + 2un - \delta_i - 1 \in I_p$ directly preceding α_i is mapped into the a-block of t at position $2pn^2 + 2un - 1$ directly preceding β . Thus, all letters of α_i are mapped into β and one has

$$\sum_{i=1}^{k} |\alpha_i| \le |\beta|. \tag{2.2}$$

By Observation 2.9(v), α_i has length $\ell_{u,(n-\delta_i/2)} = \ell_{u,d_i}$. Since β is a short b-block, it has length 2k-1. From (2.2), we get $\sum_{i=1}^k \ell_{u,d_i} \leq 2k-1$. Thus, there

is some $i^* \in \{1, ..., k\}$ with $\ell_{u, d_{i^*}} = 1$. By (2.1), that means $d_{i^*} = u$ or $\{u, d_{i^*}\}$ is an edge in G. Hence, $u \in N[D]$ and D is a dominating set of size k for of G.

It remains to prove (i) and (ii). For (i), note that $y_i(1) \ge 1$ and $y_i(x+1) \ge y_i(x) + 1$. Hence, δ_i is non-decreasing with all values being non-negative. Furthermore, for $x = 2Nn^2$, $y_i(x) \le 2Nn^2 + 2(n-1)$ since t has only so many blocks by Observation 2.9(ii). Thus, the maximum possible value of δ_i is 2(n-1). Towards (ii), we say that a value of $p \in \{0, \ldots, N-1\}$ is bad for i if δ_i is not constant over I_p . For such a p, one has $\delta_i(2pn^2+1) < \delta_i(2(p+1)n^2+1)$. Hence, there can be at most 2(n-1) values of p that are bad for i. Overall, there are at most 2k(n-1) < N values of p that are bad for some $i \in \{1, \ldots, k\}$. Thus, at least one value is not bad for any i. For this value of p, every δ_i is constant over the interval I_p .

2.3 W[2]-hardness of scheduling problems parameterized by width

In the previous section, we showed W[2]-hardness of Shuffle Product. We now transfer this result to scheduling problems on parallel identical machines.

Theorem 2.11. The following two problems are W[2]-hard parameterized by the width of the partial order giving the precedence constraints.

(i) $P2|prec, p_j \in \{1, 2\}|C_{max}$, (ii) $P3|prec, p_j = 1, size_j \in \{1, 2\}|C_{max}$.

We prove (i) using the following parameterized reduction from Shuffle Product with k+1 words to P2|prec, $p_i \in \{1,2\} | C_{\text{max}}$ with k+2 chains.

Construction 2.12. Let (s_1,\ldots,s_k,t) be a Shuffle Product instance over the alphabet $\Sigma=\{1,2\}$. Assume that $|t|_1=\sum_{i=1}^k |s_i|_1$ and $|t|_2=\sum_{i=1}^k |s_i|_2$ (otherwise, it is a no-instance). We create an instance of P2|prec, $p_j \in \{1,2\} | C_{\max}$:

- (1) For each $i \in \{1, ..., k\}$, create a chain of worker jobs $j_{i1} \prec j_{i2} \prec \cdots \prec j_{i|s_i|}$, where $j_{i,x}$ has length $s_i[x]$.
- (2) For each $x \in \{1, \ldots, |t|\}$, create three floor jobs $z_{x,1}, z_{x,2}, z_{x,3}$ with $z_{x,1} \prec z_{x,2}$ and $z_{x,1} \prec z_{x,3}$, where $z_{x,1}$ has length t[x], and $z_{x,2}$ and $z_{x,3}$ have length 1. If x < |t|, then also add the precedence constraints $z_{x,2} \prec z_{x+1,1}$ and $z_{x,3} \prec z_{x+1,1}$.

Observe that $\{z_{x,1}, z_{x,2} \mid 1 \le x \le |t|\}$ is chain. Thus, the makespan of any schedule is at least $T := \sum_{x=1}^{|t|} (t[x]+1)$. For $x \in \{1, ..., n\}$, let $\tau(x) := \sum_{i=1}^{x-1} (t[x]+1)$.

Observation 2.13. A schedule with makespan T must schedule job $z_{x,1}$ at time $\tau(x)$, and jobs $z_{x,2}$ and $z_{x,3}$ at time $\tau(x) + t[x]$. Thus, for $x \in \{1, \ldots, |t|\}$, both machines are used by floor jobs from $\tau(x) + t[x]$ to $\tau(x) + t[x] + 1$ and one machine is free of floor jobs between $\tau(x)$ and $\tau(x) + t[x]$ for t[x] time units. We call these available time slots.

Construction 2.12 runs in polynomial time. Moreover, from k+1 input words, it creates instances of width k+2: there are k chains of worker jobs and the floor decomposes into two chains $\{z_{x,1}, z_{x,2} \mid 1 \leq x \leq |t|\}$ and $\{z_{x,1}, z_{x,3} \mid 1 \leq x \leq |t|\}$. To prove Theorem 2.11(i), one can thus show that $t \in s_1 \sqcup \cdots \sqcup s_k$ if and only if the created P2|prec, $p_j \in \{1,2\} | C_{\text{max}}$ instance allows for a schedule of makespan T. By

Observation 2.13, any such schedule has available time slots of lengths corresponding to the letters in t, each of which can accommodate a worker job corresponding to a letter of s_1, \ldots, s_k . The precedence constraints ensure that these worker jobs get placed into the time slots corresponding to letters of t in increasing order.

The proof of Theorem 2.11(ii) works analogously: one simply replaces worker jobs of length two by worker jobs of length one that require two machines and modifies the floor jobs so that they do not create time slots of length one or two, but so that each created time slot is available on only one or on two machines. To achieve this, the construction uses three machines.

3 Resource-Constrained Project Scheduling

In Section 2.3, we have seen that P3|prec, $p_j=1$,size $_j \in \{1,2\}|C_{\text{max}}$ is W[2]-hard parameterized by the partial order width. It follows that also RCPSP (cf. Problem 1.1) is W[2]-hard for this parameter, even if the number of resources and the maximal resource usage are bounded by two and all jobs have unit processing times. In this section, we additionally consider the lag parameter:

Definition 3.1 (earliest possible starting time, lag). Let $J_0 \subseteq J$ be the jobs that are minimal elements in the partial order \preceq . The *earliest possible starting time* σ_j is 0 for a job $j \in J_0$ and, inductively, $\max_{i \neq j} (\sigma_i + p_i)$ for a job $j \in J \setminus J_0$. The *lag* of a feasible schedule $(s_i)_{i \in J}$ is $\lambda := \max_{i \in J} s_i - \sigma_i$.

Lenstra and Rinnooy Kan's NP-hardness proof for P|prec, $p_j=1|C_{\max}$ [10] shows that it is even NP-hard to decide whether there is a schedule of makespan at most three and lag at most one. Thus, the lag λ alone cannot lead to a fixed-parameter algorithm for RCPSP, just as the width w alone cannot. We show a fixed-parameter algorithm for the parameter $\lambda + w$.

Theorem 3.2. An optimal schedule with lag at most λ for RCPSP is computable in $(2\lambda + 1)^w \cdot 2^w \cdot \text{poly}(n)$ time if it exists, where w is the partial order width.

Our algorithm is a refinement of Servakh's pseudo-polynomial-time algorithm for RCPSP with constant width [14], which is based on graphical optimization methods introduced by Akers [1] and Hardgrave and Nemhauser [9] for hand-optimizing Job Shop schedules for two jobs. We provide a concise translation of Servakh's algorithm in Section 3.1 before we prove Theorem 3.2 in Section 3.2.

3.1 Geometric interpretation of RCPSP

Given an RCPSP instance with precedence constraints \leq of width w, by Dilworth's theorem, we can decompose our set J of jobs into w pairwise disjoint chains. More specifically, these chains are efficiently computable [7]. For $\ell \in \{1, \ldots, w\}$, denote the jobs in chain ℓ by a sequence $(j_{\ell k})_{k=1}^{n_{\ell}}$ such that $j_{\ell k} \prec j_{\ell k+1}$ and let

 $L^i_\ell := \sum_{k=1}^i p_{j_{\ell k}}$ be the sum of processing times of the first i jobs on chain ℓ , $L_\ell := L^{n_\ell}_\ell$ be the sum of processing times of all jobs on chain ℓ .

Let $\mathbf{0} := (0, \dots, 0) \in \mathbb{R}^w$ and $\mathbf{L} := (L_1, \dots, L_w)$. Each point in the w-dimensional orthotope $X := \{ \mathbf{x} \in \mathbb{R}^w \mid \mathbf{0} \le \mathbf{x} \le \mathbf{L} \}$ describes a *state* as follows.

Definition 3.3 (running, completed, feasibility). Let $\boldsymbol{x} = (x_1, \dots, x_w) \in X$. For each chain $\ell \in \{1, \dots, w\}$, if $x_\ell \in [L_\ell^{i-1}, L_\ell^i)$, then the jobs $(j_{\ell k})_{k=1}^{i-1}$ of chain ℓ are completed and job $j_{\ell i}$ has been processed for $x_\ell - L_\ell^{i-1}$ time. We call job $j_{\ell i}$ running if $L_\ell^{i-1} < x_\ell < L_\ell^i$. We denote by

 $J(x) \subseteq J$ the set of jobs running in state x and by $C(x) \subseteq J$ the set of jobs completed in state x.

A point $x \in X$ is feasible if it holds that both

- (IF1) the jobs $J(\boldsymbol{x})$ comply with resource constraints, that is, $\sum_{j \in J(\boldsymbol{x})} r_{j\rho} \leq R_{\rho}$ for each resource $\rho \in R$, and
- (IF2) if there are two jobs $i \prec j$ such that $j \in J(x)$, then $i \in C(x)$.

Note that points $x \in X$ may indeed violate (IF2): there are not only precedence constraints between jobs on one chain, but also between jobs on different chains.

Each feasible schedule now yields a path of feasible points in the orthotope X from the point $\mathbf{0}$, where no job has started, to the point \mathbf{L} , where all jobs are completed. Each such path consists of (linear) segments of the form $[\boldsymbol{x}, \boldsymbol{x} + t\boldsymbol{\delta}]$ for some $\boldsymbol{\delta} = (\delta_1, \dots, \delta_w) \in \{0, 1\}^w$, which corresponds to running exactly the jobs on the chains ℓ with $\delta_{\ell} = 1$ for t units of time. Since all processing times and starting times are integers (cf. Problem 1.1), we can assume $t \in \mathbb{N}$.

Definition 3.4 (feasibility of segments and their lengths). The *length* of a segment $[x, x + t\delta]$ is t. The *length* of a path is the sum of the lengths of its segments. A segment $[x, x + t\delta]$ is *feasible* if it contains only feasible points and interrupts no jobs; that is, if there is a job $j \in J(x)$ on chain ℓ , then $\delta_{\ell} = 1$.

There is now a one-to-one correspondence between feasible schedules and paths from $\bf 0$ to $\bf L$ consisting only of feasible segments and between the shortest of these paths and optimal schedules. This leads to the following algorithm.

Algorithm 3.5 (Servakh [14]). Compute a shortest feasible path from $\mathbf{0}$ to L using dynamic programming: for each feasible point $\mathbf{x} \in X \cap \mathbb{N}^w$ in lexicographically increasing order, compute the length $P(\mathbf{x})$ of a shortest feasible path from $\mathbf{0}$ to \mathbf{x} using the recurrence relation

$$P(\mathbf{0}) = 0$$
, $P(x) = \min_{\delta \in \Delta_x} P(x - \delta) + 1$ for feasible $x \in X \cap \mathbb{N}^w \setminus \{\mathbf{0}\}$, (3.1)

where Δ_x is the set of vectors $\delta \in \{0,1\}^w$ such that segment $[x-\delta,x]$ is feasible.

To compute $P(\mathbf{L})$, one thus iterates over at most $\prod_{\ell=1}^w (L_\ell + 1)$ points $\mathbf{x} \in X \cap \mathbb{N}^w$, for each of them over 2^w vectors $\mathbf{\delta} \in \{0,1\}^w$, and, for each, decides whether $[\mathbf{x} - \mathbf{\delta}, \mathbf{x}]$ is feasible. Since the set of running jobs is the same for all interior points of the segment, it is enough to check the feasibility of its end points and one interior point, which can be done in polynomial time. Thus, the algorithm runs in $\prod_{\ell=1}^w (L_\ell + 1) \cdot 2^w \cdot \operatorname{poly}(n)$ time, which is pseudo-polynomial for constant w.

3.2 Fixed-parameter algorithm for arbitrary processing times

The bottleneck of Algorithm 3.5 is that it searches for a shortest path from $\mathbf{0}$ to \mathbf{L} in the whole orthotope X. For the case where we are only accepting schedules of maximum lag λ , we will shrink the search space significantly: we show that we only have to search for paths within a tight corridor around the path corresponding to the schedule $(\sigma_i)_{i \in J}$ that starts jobs at the earliest possible time.

Definition 3.6 (point at time t **on a path).** Let p be the path from **0** to L corresponding to a not necessarily feasible schedule $(s_j)_{j\in J}$ that, however, respects precedence constraints. Let $t \geq 0$ and T be the length of p.

Then, p(t) is the endpoint of the subpath of length t of p starting in $\mathbf{0}$ for $t \leq T$, and $p(t) := \mathbf{L}$ for t > T.

Since the definition requires $(s_j)_{j\in J}$ to respect precedence constraints, p(t) determines the state (cf. Definition 3.3) at time t according to schedule $(s_j)_{j\in J}$.

Definition 3.7 (λ -corridored). Let p be the path corresponding to the schedule $(\sigma_j)_{j\in J}$ that starts jobs at the earliest possible time (cf. Definition 3.1).

$$\Gamma_{\lambda}(t) := \{ \boldsymbol{x} \in X \mid \boldsymbol{p}(t) - \boldsymbol{\lambda} \leq \boldsymbol{x} \leq \boldsymbol{p}(t) \}, \text{ where } \boldsymbol{\lambda} = (\lambda, \dots, \lambda) \in \mathbb{N}^{w}.$$

We call a path $q \lambda$ -corridored if $\mathbf{q}(t) \in \Gamma_{\lambda}(t)$ for all $t \geq 0$.

Note that points on the path p in Definition 3.7 may violate Definition 3.3(IF1), but not (IF2). One can show the following relation between λ -corridored paths and schedules of lag λ .

Lemma 3.8. A feasible schedule $(s_j)_{j\in J}$ has lag at most λ if and only if its corresponding path q is λ -corridored.

Lemma 3.8 allows us to compute a shortest feasible path from $\mathbf{0}$ to \mathbf{L} using only points in $\Gamma_{\lambda}(t)$ for some t. Herein, we will exploit the following condition for checking whether a path segment can be part of a λ -corridored path.

Lemma 3.9. Let $[\boldsymbol{x}, \boldsymbol{x} + t\boldsymbol{\delta}]$ for $\boldsymbol{\delta} \in \{0, 1\}^w$. If $\boldsymbol{x} \in \Gamma_{\lambda}(t_0)$ and $\boldsymbol{x} + t\boldsymbol{\delta} \in \Gamma_{\lambda}(t_0 + t)$ for some $t_0 \geq 0$, then $\boldsymbol{x} + \tau\boldsymbol{\delta} \in \Gamma_{\lambda}(t_0 + \tau)$ for all $0 \leq \tau \leq t$.

Proof. Let p be the path corresponding to schedule $(\sigma_j)_{j\in J}$ as in Definition 3.7 and let $\boldsymbol{\delta}=(\delta_1,\ldots,\delta_w)\in\{0,1\}^w$. For any $\tau\in[0,t]$, consider

$$\boldsymbol{x}^{\tau} = (x_1^{\tau}, \dots, x_w^{\tau}) := \boldsymbol{x} + \tau \boldsymbol{\delta}$$
 and $\boldsymbol{y}^{\tau} = (y_1^{\tau}, \dots, y_{\ell}^{\tau}) := \boldsymbol{p}(t_0 + \tau).$

By the prerequisites of the lemma, we have $y^0 - \lambda \le x^0 \le y^0$ and $y^t - \lambda \le x^t \le y^t$. We show $y^\tau - \lambda \le x^\tau \le y^\tau$ for any $\tau \in [0, t]$.

We start with $\boldsymbol{x}^{\tau} \leq \boldsymbol{y}^{\tau}$. For the sake of contradiction, assume that there is some chain ℓ and a $\tau \in [0,t]$ such that $x_{\ell}^{\tau} > y_{\ell}^{\tau}$. Then, $x_{\ell}^{\tau} > y_{\ell}^{\tau} \geq y_{\ell}^{0} \geq x_{\ell}^{0}$. It follows that $\delta_{\ell} = 1$, which contradicts $\boldsymbol{x}^{t} \leq \boldsymbol{y}^{t}$ because, then,

$$x_{\ell}^{t} = x_{\ell}^{0} + t = x_{\ell}^{0} + \tau + (t - \tau) = x_{\ell}^{\tau} + (t - \tau) > y_{\ell}^{\tau} + (t - \tau) \ge y_{\ell}^{\tau + (t - \tau)} = y_{\ell}^{t}.$$

Now, we show $\boldsymbol{y}^{\tau} - \boldsymbol{\lambda} \leq \boldsymbol{x}^{\tau}$. Consider some chain ℓ . If $\delta_{\ell} = 1$, then we have $y_{\ell}^{\tau} - \lambda \leq y_{\ell}^{0} + \tau - \lambda \leq x_{\ell}^{0} + \tau = x_{\ell}^{\tau}$ and we are fine. If $\delta_{\ell} = 0$ and there is a $\tau \in [0, t]$ such that $y_{\ell}^{\tau} - \lambda > x_{\ell}^{\tau}$, then $y_{\ell}^{t} - \lambda \geq y_{\ell}^{\tau} - \lambda > x_{\ell}^{\tau} = x_{\ell}^{t}$, contradicting $\boldsymbol{y}^{t} - \boldsymbol{\lambda} \leq \boldsymbol{x}^{t}$. \square

We can now prove the following result by computing recurrence (3.1) for each of the $(\lambda + 1)^w$ feasible points $\boldsymbol{x} \in \Gamma_{\lambda}(t) \cap \mathbb{Z}^w$ for all $t \in \{0, \dots, L\}$.

Proposition 3.10. An optimal schedule of lag at most λ for RCPSP if it exists is computable in $(\lambda + 1)^w \cdot 2^w \cdot \text{poly}(L)$ time, where L is the sum of all processing times and w is the partial order width.

However, note that this is a fixed-parameter algorithm only for polynomial processing times, which is why we skip the proof and go on towards proving Theorem 3.2—a fixed-parameter algorithm that works for arbitrarily large processing times. To this end, we prove that all maximal segments of a path corresponding to a schedule with lag at most λ start and end in one of $2 \cdot |J|$ hypercubes with edge length $2\lambda + 1$.

Lemma 3.11. Let q be the path of a feasible schedule $(s_j)_{j \in S}$ of lag at most λ and let $t_2 \leq t_1 \leq t_2 + \lambda$. Then, $q(t_1) \in \Gamma_{2\lambda}(t_2 + \lambda)$ (cf. Definition 3.7).

Proof. Consider the schedule $(\sigma_j)_{j\in J}$ that starts each job at the earliest possible time and its path p. Our aim is to show

$$p(t_2 + \lambda) - 2\lambda \le q(t_1) \le p(t_2 + \lambda),$$

where $\lambda = (\lambda, ..., \lambda) \in \mathbb{N}^w$. By Lemma 3.8, q is λ -corridored. Thus,

$$p(t_1) - \lambda \le q(t_1) \le p(t_1)$$
 and $p(t_2 + \lambda) - \lambda \le q(t_2 + \lambda) \le p(t_2 + \lambda)$.

From this, one easily gets $q(t_1) \leq p(t_1) \leq p(t_2 + \lambda)$. Moreover, one has

$$p(t_2 + \lambda) - 2\lambda \le q(t_2 + \lambda) - \lambda \le q(t_2) + \lambda - \lambda = q(t_2) \le q(t_1).$$

Lemma 3.12. Let q be the path of a feasible schedule $(s_j)_{j\in S}$ of lag at most λ and let $[x, x + t\delta]$ be a maximal segment of q such that the set $J(x + \tau\delta)$ of running jobs (cf. Definition 3.3) is the same for all $\tau \in (0, t)$. Then,

$$\{x, x + t\delta\} \subseteq \Gamma := \bigcup_{j \in J} \Gamma_{2\lambda}(\sigma_j + \lambda) \cup \bigcup_{j \in J} \Gamma_{2\lambda}(\sigma_j + p_j + \lambda),$$

where $(\sigma_j)_{j\in J}$ is the schedule that starts each job at the earliest possible time.

Proof. Let t_0 be chosen arbitrarily such that $q(t_0) \in \{x, x + t\delta\}$. By maximality of the segment, some job $j \in J$ is starting or ending at time t_0 , that is, $t_0 = s_j$ or $t_0 = s_j + p_j$. Then, $\{x, x + t\delta\} \subseteq \Gamma_{2\lambda}(\sigma_j + \lambda) \cup \Gamma_{2\lambda}(\sigma_j + p_j + \lambda)$ follows from $\sigma_j \leq s_j \leq \sigma_j + \lambda$ and Lemma 3.11.

We are now ready to show a fixed-parameter algorithm for RCPSP parameterized by length and maximum lag. That is, we prove Theorem 3.2.

Proof (of Theorem 3.2). We compute the shortest feasible λ -corridored path from the state $\mathbf{0}$, were no job has started, to the state \mathbf{L} , where all jobs have been completed (cf. Lemma 3.8). We use dynamic programming similarly to Algorithm 3.5. By Lemma 3.12, it is enough to consider those paths whose segments start and end in Γ . Thus, for each $\mathbf{x} \in \Gamma \cap \mathbb{N}^w$ in lexicographically increasing order, we compute the length $P(\mathbf{x})$ of a shortest λ -corridored path from $\mathbf{0}$ to \mathbf{x} with segments starting and ending in Γ . To this end, for an $\mathbf{x} \in \Gamma \cap \mathbb{N}^w$, let $\Delta_{\mathbf{x}}$ be the set of vectors $\mathbf{\delta} \in \{0,1\}^w$ such that,

- (i) there is a smallest integer $t_{\delta} \geq 1$ such that $x t_{\delta} \cdot \delta \in \Gamma$ and such that
- (ii) the segment $[\boldsymbol{x} t_{\boldsymbol{\delta}} \cdot \boldsymbol{\delta}, \boldsymbol{x}]$ is feasible.

Then, $P(\mathbf{0}) = 0$ and, for feasible $x \in \Gamma \cap \mathbb{N}^w \setminus \{\mathbf{0}\}$, one has

$$P(x) = \min\{P(x - t_{\delta} \cdot \delta) + t_{\delta} \mid \delta \in \Delta_x \text{ and } x \in \Gamma_{\lambda}(P(x - t_{\delta} \cdot \delta) + t_{\delta})\},$$

where $\min \emptyset = \infty$ and the last condition on \boldsymbol{x} uses Lemma 3.9 to ensure that we are indeed computing the length $P(\boldsymbol{x})$ of a λ -corridored path (cf. Definition 3.7) to \boldsymbol{x} : by induction, we know that $P(\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta})$ is the length of a shortest λ -corridored path to $\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta}$, and thus $\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta} \in \Gamma_{\lambda}(P(\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta}))$.

We have to discuss how to check (i) and (ii). One can check (ii) in polynomial time since it is enough to check feasibility at the end points and one interior point of the segment since the set of jobs running at the interior points of $[\boldsymbol{x}-t_{\boldsymbol{\delta}}\cdot\boldsymbol{\delta},\boldsymbol{x}]$ does not change: otherwise, since jobs are started or finished only at integer times, there is a maximal subsegment $[\boldsymbol{x},\boldsymbol{x}-t\cdot\boldsymbol{\delta}]$ with $t\leq t_{\boldsymbol{\delta}}-1$ where the set of running jobs does not change. Then $\boldsymbol{x}-t\cdot\boldsymbol{\delta}\in \Gamma$ by Lemma 3.12, contradicting the minimality of $t_{\boldsymbol{\delta}}$.

Towards (i), we search for the minimum $t_{\delta} \geq 1$ such that $\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta} \in \Gamma$. Consider the schedule $(\sigma_j)_{j \in J}$ that schedules each job at the earliest possible time (cf. Definition 3.1). It is computable in polynomial time. By Lemma 3.12, we search for the minimum $t_{\delta} \geq 1$ such that $\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta} \in \Gamma_{2\lambda}(\sigma_j + \lambda)$ or $\boldsymbol{x} - t_{\delta} \cdot \boldsymbol{\delta} \in \Gamma_{2\lambda}(\sigma_j + \lambda)$ for some job $j \in J$. That is, by Definition 3.7, for each job j, we find the minimum $t_j \geq 1$ that solves a system of linear inequalities of the form $\boldsymbol{y} - 2\boldsymbol{\lambda} \leq \boldsymbol{x} - t_j \cdot \boldsymbol{\delta} \leq \boldsymbol{y}$, where $\boldsymbol{\delta} = (\delta_1, \dots, \delta_w) \in \{0, 1\}^w$. Writing $\boldsymbol{y} = (y_1, \dots, y_w)$ and $\boldsymbol{x} = (x_1, \dots, x_w)$, either $t_j = \max(\{1\} \cup \{x_\ell - y_\ell \mid \delta_\ell = 1\})$ is the minimum such t_j or there is no solution for job j. Note that t_j is an integer since \boldsymbol{x} and \boldsymbol{y} are integer vectors. Thus, $t_{\delta} = \min_{j \in J} t_j$ is computable in polynomial time.

We conclude that we process each $\boldsymbol{x} \in \Gamma \cap \mathbb{N}^w$ in $2^w \cdot \operatorname{poly}(n)$ time. Moreover, Γ contains at most $2 \cdot |J| \cdot (2\lambda + 1)^w$ integer points since each job $j \in J$ contributes at most $(2\lambda + 1)^w$ points in $\Gamma_{2\lambda}(\sigma_j + \lambda)$ and at most $(2\lambda + 1)^w$ points in $\Gamma_{2\lambda}(\sigma_j + p_j + \lambda)$. A total running time of $(2\lambda + 1)^w \cdot 2^w \cdot \operatorname{poly}(n)$ follows. \square

4 Conclusion

Our algorithm for RCPSP shows, in particular, that P3|prec, p_j =1| C_{\max} is fixed-parameter tractable parameterized by the partial order width w and allowed lag λ . Since the NP-hardness of this problem is a long-standing open question [8, OPEN8], it would be surprising to show W[1]-hardness of this problem for any parameter: this would exclude polynomial-time solvability unless FPT = W[1]. Thus, it makes sense to search for a fixed-parameter algorithm for P3|prec, p_j =1| C_{\max} parameterized by w, whereas we showed that already P2|prec, p_j ={1,2}| C_{\max} and P3|prec, p_j =1,size $_j$ ={1,2}| C_{\max} are W[2]-hard parameterized by w.

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A Appendix: Omitted proofs

A.1 Proofs for Section 2.1

Theorem 2.1. P2|chains| C_{max} is weakly NP-hard even for precedence constraints consisting of three chains.

Proof. We reduce from the weakly NP-hard Partition problem [8, SP12]: Given a multiset of positive integers $A = \{a_1, \ldots, a_t\}$, decide whether there is a subset $A' \subseteq A$ such that $\sum_{a_i \in A'} a_i = \sum_{a_i \in A \setminus A'} a_i$. Let $A = \{a_1, \ldots, a_t\}$ be a Partition instance. If $b := \left(\sum_{a_i \in A} a_i\right)/2$ is not an integer, then we are facing a no-instance. Otherwise, we construct a P2|chains| C_{\max} instance as follows. Create three chains of jobs $J^0 := \{j_1^0 \prec \cdots \prec j_t^0\}$, $J^1 := \{j_1^1 \prec \cdots \prec j_{t+1}^1\}$, and $J^2 := \{j_1^2 \prec \cdots \prec j_{t+1}^2\}$. The jobs j_i^0 with $i \in \{1, \ldots, t\}$ have processing time a_i . The jobs j_i^0 with $\ell \in \{1, 2\}$ and $\ell \in \{1, \ldots, t+1\}$ have processing time $\ell \in \{1, 2\}$ and $\ell \in \{1, \ldots, t+1\}$ have processing time $\ell \in \{1, 2\}$ and the input Partition instance is a yes-instance of and only if the created P2|chains| $\ell \in \{0, 1\}$ instance allows for a schedule with makespan $\ell \in \{1, 2\}$ instance and partition instance allows for a schedule with makespan $\ell \in \{1, 2, 3\}$.

 (\Leftarrow) Assume that our constructed P2|chains| C_{max} instance has a schedule of makespan T. Since we have two chains J^1 and J^2 each containing t+1 long jobs with processing time 2b, each machine must perform exactly t+1 long jobs from $J^1 \cup J^2$ and may perform additional short jobs with processing times at most b from J^0 . Let A' be the set of elements in A corresponding to the jobs from J^0 processed by the first machine. Then, $A \setminus A'$ corresponds to the jobs from J^0 that are performed by the second machine. Since the makespan is T = (2t + 3)band the long jobs already need (2t+2)b time units on each machine, it holds that $\sum_{a_i \in A'} a_i = \sum_{a_i \in A \setminus A'} a_i = b$. Thus, A' is a solution for our Partition instance. (\Rightarrow) Let $A' \subseteq A$ with $\sum_{a_i \in A'} a_i = \sum_{a_i \in A \setminus A'} a_i = b$ be a solution for PARTI-TION and let $J^{\prime 0} \subseteq J^0$ denote the set of jobs corresponding to the elements in $A' \subseteq A$. We construct a schedule of makespan T as follows. First, ignoring the jobs in J^0 , schedule each jobs in $J^1 \cup J^2$ at the earliest possible time, that is, the starting time of each job equals the sum of processing times of all preceding jobs in its chain. So far, the maximum completion time is (2t+2)b and each chain is processed by one machine. Next, we modify this schedule by "inserting" the jobs from J^0 in between two already scheduled jobs. Herein, inserting job j between job $j', j'' \in J^z$ for $z \in \{1,2\}$ means to set the starting time of j to the starting time of j'' and to increase the starting times of j'' and of all its successors in J^z by the processing time of j. We insert all jobs from J_0 according to the precedence constraints on J_0 : we insert job j_i^0 between jobs j_i^1 and j_{i+1}^1 if $a_i \in A'$ and between job j_i^2 and job j_{i+1}^2 if $a_i \in A \setminus A'$. The way we defined the insertion operation ensures that the schedule can still realized by two machines. Since $\sum_{a_i \in A'} a_i = \sum_{a_i \in A \setminus A'} a_i = b$, it further holds that the starting time of every job was increased by at most b and, hence, the constructed schedule has makespan at most T = (2t+3)b (the latest executed jobs are still job j_{t+1}^1 and job j_{t+1}^2). It remains to check that the precedence constraints are fulfilled: constraints between jobs from J^1 or from J^2 remain fulfilled since we do not change their relative execution order. The constraints between

jobs j_i^0 and $j_{i'}^0$ for $1 \le i < i' \le t$ are fulfilled because, in our constructed schedule, the starting time of job j_i^0 is in the interval $[i \cdot 2b, i \cdot 2b + b)$, the starting time of job $j_{i'}^0$ is in the interval $[i' \cdot 2b, i' \cdot 2b + b)$, and these interval do not intersect. \square

A.2 Proofs of Section 2.3

Lemma A.2. A SHUFFLE PRODUCT instance $(t, s_1, ..., s_k)$ is a yes-instance if and only if the P2|prec, $p_j \in \{1, 2\} | C_{\text{max}}$ instance created by Construction 2.12 allows for a schedule of makespan $T := \sum_{r=1}^{|t|} (t[x] + 1)$.

Proof. (\Rightarrow) Consider k functions f_1, \ldots, f_k mapping the letters of s_1, \ldots, s_k to the letters of t as required by Definition 2.2. We use the schedule described by Observation 2.13 for the floor jobs, and, for $i \in \{1, \ldots, k\}$ and $x \in \{1, \ldots, |s_i|\}$, we schedule the worker job $j_{i,x}$ to time $\tau(f_i(x))$. Note that the precedence constraints of the worker jobs are satisfied since the functions f_i are strictly increasing and the difference between two consecutive values of τ is at least 2 (which is the maximal length of a job). Moreover, for each $y \in \{1, \ldots, |t|\}$ there is exactly one i such that $y = f_i(x)$ for some x. Hence, a worker job $j_{i,x}$ can use the available time slot at $\tau(y)$ without any other worker job occupying it. Finally, this worker job $j_{i,x}$ needs time $s_i[x] = t[f_i(x)] = t[y]$, which is exactly the length of the available time slot at $\tau(y)$.

(\Leftarrow) Consider a scheduling with makespan T, and let $n_1 = |t|_1$ and $n_2 = |t|_2$. We construct functions f_1, \ldots, f_k mapping the letters of s_1, \ldots, s_k to the letters of t as required by Definition 2.2. From Observation 2.13, we know that only time slots of lengths one and two are available for worker jobs. Hence, each job $j_{i,x}$ with length $s_i[x] = 2$ is scheduled to a time $\tau(y)$ for some $y \in \{1, \ldots, |t|\}$ with t[y] = 2. We put $f_i(x) := y$. Since the number of worker jobs of length two is $\sum_{i=1}^{k} |s_i|_2 = n_2$, all available time slots of length two are used by worker jobs of length two. Thus, jobs of length one must use pairwise distinct available time slots of length one. Thus, each job $j_{i,x}$ of length $s_i[x] = 1$ is scheduled to a time $\tau(y)$ for some $y \in \{1, \dots, |t|\}$ with t[y] = 1. Put $f_i(x) := y$. Clearly, each of the constructed functions f_i is total on $\{1,\ldots,|s_i|\}$. Moreover, no two functions share a value since the execution times of worker jobs do not intersect. Finally, each function is increasing: due to the precedence constraints, job $j_{i,x}$ is scheduled before $j_{i,x+1}$ and thus, $\tau(f_i(x)) < \tau(f_i(x+1))$ and $f_i(x) < f_i(x+1)$ since τ is increasing. Finally, note that $t[f_i(x)] = s_i[x]$ by construction, hence functions f_i define a valid mapping of the letters of s_i into t.

A.3 Proofs of Section 3.2

Lemma 3.8. A feasible schedule $(s_j)_{j\in J}$ has lag at most λ if and only if its corresponding path q is λ -corridored.

Proof. (\Rightarrow) Consider a point q(t) on q and the corresponding point p(t) on the path p corresponding to schedule $(\sigma_j)_{j\in J}$ (cf. Definition 3.7). Then one has $q(t) \leq p(t)$ since, at time t, schedule $(s_j)_{j\in J}$ cannot have processed any chain

for more time than schedule $(\sigma_j)_{j\in J}$. Moreover, one has $\boldsymbol{q}(t)\geq \boldsymbol{p}(t)-\boldsymbol{\lambda}$ since, at time t, a chain ℓ that has been processed for x_ℓ time by schedule $(\sigma_j)_{j\in J}$ has been processed for at least $x_\ell-\lambda$ time by schedule $(s_j)_{j\in J}$. Thus, $\boldsymbol{q}(t)\in \Gamma_\lambda(t)$. (\Leftarrow) We show that $s_j-\sigma_j\leq \lambda$ for an arbitrary job j. Since $\boldsymbol{q}(s_j)\in \Gamma_\lambda(s_j)$, one has $\boldsymbol{p}(s_j)-\boldsymbol{\lambda}\leq \boldsymbol{q}(s_j)\leq \boldsymbol{p}(s_j)$. In particular $\boldsymbol{q}(s_j)\leq \boldsymbol{p}(s_j)+\boldsymbol{\lambda}$. In state $\boldsymbol{q}(s_j)$, job j has not been processed for any time yet. It follows that it has been processed for at most λ time in state $\boldsymbol{p}(s_j)$. Since this is the state of schedule $(\sigma_j)_{j\in J}$ at time s_j , we get $\sigma_j\geq s_j-\lambda$, that is, $s_j-\sigma_j\leq \lambda$.