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Flip Distance Between Triangulations of a Simple Polygon is NP-Complete

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Abstract Let T be a triangulation of a simple polygon. A *flip* in T is the operation of replacing one diagonal of T by a different one such that the resulting graph is again a triangulation. The *flip distance* between two triangulations is the smallest number of flips required to transform one triangulation into the other. For the special case of convex polygons, the problem of determining the shortest flip distance between two triangulations is equivalent to determining the rotation distance between two binary trees, a central problem which is still open after over 25 years of intensive study.

We show that computing the flip distance between two triangulations of a simple polygon is NP-hard. This complements a recent result that shows APX-hardness of determining the flip distance between two triangulations of a planar point set.

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1 Introduction

Let P be a simple polygon in the plane, that is, a closed region bounded by a piecewise linear, simple cycle. A *triangulation* of P is a geometric (straight-line) maximal outerplanar graph whose outer face is the complement of P and whose vertex set consists of the vertices of P. The edges that are not on the outer face are called *diagonals*. Let d be a diagonal whose removal creates a convex quadrilateral. Replacing d with the other diagonal of the quadrilateral yields another triangulation of P. This operation is called a *flip*. The *flip graph* of P is the abstract graph whose vertices are the triangulations of P and in which two triangulations are adjacent if and only if they differ by a single flip. We study the *flip distance*, i.e., the minimum number of flips required to transform a given source triangulation into a target triangulation.

Edge flips became popular in the context of Delaunay triangulations. Lawson [15] proved that any triangulation of a planar n-point set can be transformed into any other by $O(n^2)$ flips. Hence, for every planar n-point set the flip graph is connected with diameter $O(n^2)$. Later, Lawson showed that in fact every triangulation can be transformed to the Delaunay triangulation by $O(n^2)$ flips that locally fix the Delaunay property [16]. Hurtado, Noy, and Urrutia [11] gave an example where the flip distance is $\Omega(n^2)$, and they showed that the same bounds hold for triangulations of simple polygons. They also proved that if the polygon has k reflex vertices, then the flip graph has diameter $O(n+k^2)$. In particular, the flip graph of any planar polygon has diameter $O(n^2)$. Their result also generalizes the well-known fact that the flip distance between any two triangulations of a convex polygon is at most $2n-10$, for $n > 12$. This was shown by Sleator, Tarjan, and Thurston [22] in their work on the flip distance in convex polygons. The latter case is particularly interesting due to the correspondence between flips in triangulations of convex polygons and rotations in binary trees: The dual graph of such a triangulation is a binary tree, and a flip corresponds to a rotation in that tree; conversely, for every binary tree, a triangulation can be constructed.

We mention two further remarkable results on flip graphs for point sets. Hanke, Ottmann, and Schuierer [10] showed that the flip distance between two triangulations is

bounded by the number of crossings in their overlay. Eppstein [9] gave a polynomial-time algorithm for calculating a lower bound on the flip distance. His bound is tight for point sets with no empty 5-gons; however, except for small instances, such point sets are not in general position (i.e., they must contain collinear triples) [1]. A recent survey on flips is provided by Bose and Hurtado [4].

Recently, the problem of finding the flip distance between two triangulations of a point set was shown to be NP-hard by Lubiw and Pathak [18] and, independently, by Pilz [19]. The latter proof was later improved to show APX-hardness of the problem. A recent preprint shows that the problem is fixed-parameter tractable [14]. Here, we show that the corresponding problem remains NP-hard even for simple polygons. This can be seen as a further step towards settling the complexity of deciding the flip distance between triangulations of convex polygons or, equivalently, the rotation distance between binary trees. This variant of the problem was probably first addressed by Culik and Wood [7] in 1982 (showing a flip distance of $2n-6$) in the context of similarity measures between trees.

We now give the formal problem definition: given a simple polygon P, two triangulations T_1 and T_2 of P, and an integer l, decide whether T_1 can be transformed into T_2 by at most l flips. We call this decision problem PolyFlip. To show NP-hardness, we give a polynomial-time reduction from the problem Rectilinear Steiner Arborescence to PolyFlip. Rectilinear Steiner Arborescence was shown to be NP-hard by Shi and Su [21]. In Section 2, we describe the problem in detail. We present the well-known *double chain* (used by Hurtado, Noy, and Urrutia [11] for giving their lower bound), a major building block in our reduction, in Section 3. Finally, in Section 4, we describe our reduction and prove that it is correct.

2 The Rectilinear Steiner Arborescence Problem

Let S be a set of N points in the plane whose coordinates are nonnegative integers. The points in S are called *sinks*. A *rectilinear tree* A is a connected acyclic collection of horizontal and vertical line segments that intersect only at their endpoints. The *length* of A is the total length of all segments in A (cf. [13, p. 205]). The tree A is a *rectilinear Steiner tree* for S if every sink in S appears as an endpoint of a segment in A. We call A a *rectilinear Steiner arborescence* (RSA) for S if (i) A is rooted at the origin; (ii) every

leaf of A lies at a sink in S; and (iii) for each $s = (x_s, y_s) \in S$, the length of the path in A from the origin to s equals $x_s + y_s$, i.e., all edges in A point north or east, as seen from the origin [20]. In the problem Rectilinear Steiner Arborescence, we are given a set of sinks S and an integer k. The question is whether there is an RSA for S of length at most k. Shi and Su showed that Rectilinear Steiner Arborescence is strongly NP-complete; in particular, it remains NP-complete if S is contained in an $n \times n$ grid, with n polynomially bounded in N, the number of sinks [21].[1]

We will need the following important structural property of the RSA. Let A be an RSA for a set S of sinks. Let e be a vertical segment in A that does not contain a sink. Suppose there is a horizontal segment f incident to the upper endpoint a of e. Since A is an arborescence, a is the left endpoint of f. Suppose further that a is not the lower endpoint of another vertical edge. Take a copy e' of e and translate it to the right until e' hits a sink or another segment endpoint (this will certainly happen at the right endpoint of f); see Fig. 1. The segments e and e' define a rectangle R. The upper and left side of R are completely covered by e and (a part of) f. Since a has only two incident segments, every sink-root path in A that goes through e or f contains these two sides of R, entering the boundary of R at the upper right corner d and leaving it at the lower left corner b. We reroute every such path at d to continue clockwise along the boundary of R until it meets A again (this certainly happens at b), and we delete e and the part of f on R. In the resulting tree we subsequently remove all unnecessary segments (this happens if there are no more root-sink paths through b) to obtain another RSA A' for S. Then A' is not longer than A. This operation is called *sliding e to the right*. If similar conditions apply to a horizontal edge, we can *slide it upwards*. The *Hanan grid* for a point set is the set of all vertical and horizontal lines through its points. Through repeated segment slides in a shortest RSA, one can obtain the following theorem.

Theorem 2.1 ([20]) *Let S be a set of sinks. There is a minimum-length RSA A for S such that all segments of A are on the Hanan grid for $S \cup \{(0,0)\}$.* □

We use a restricted version of Rectilinear Steiner Arborescence, called YRSA. An instance (S, k) of YRSA differs from an instance for Rectilinear Steiner Arborescence in that we require that no two sinks in S have the same y-coordinate.

Theorem 2.2 YRSA *is strongly* NP*-complete.*

[1] Although a polynomial-time algorithm was claimed [23], it has later been shown to be incorrect [20].

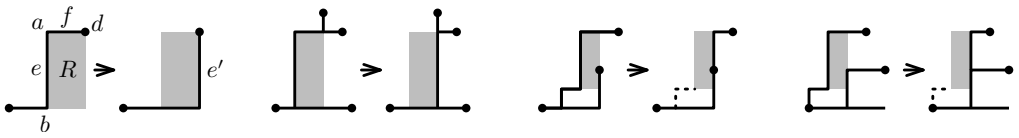

Fig. 1 The slide operation. The dots depict sinks; the rectangle R is drawn gray. The dotted segments are deleted, since they do no longer lead to a sink.

Proof Due to Theorem 2.1, YRSA and Rectilinear Steiner Arborescence are in NP [21]. We now show how to transform an instance (S, k) of Rectilinear Steiner Arborescence to an instance of YRSA. We may assume that $N = |S| \geq 3$, and we number the sinks as $S = \langle s_1, s_2, \ldots, s_N \rangle$ in an arbitrary fashion. For $i = 1, \ldots, N$, let (x_i, y_i) be the coordinates of s_i and define $s'_i := (x_i N^4, y_i N^4 + i)$. We set $S' := \{s'_1, s'_2, \ldots, s'_N\}$. The y-coordinates of the sinks in S' are pairwise distinct. We will show that there is an RSA for S of length at most k if and only if there is an RSA for S' of length at most $kN^4 + N^3$.

Let A be a rectilinear Steiner arborescence for S of length at most k. We scale A by N^4 and draw a vertical segment from each leaf to the sink in S' above it. This gives an RSA for S' of length at most $kN^4 + N^2 < kN^4 + N^3$.

Conversely, let A' be an RSA for S' of length at most $kN^4 + N^3$. Due to Theorem 2.1, we can assume that A' is on the Hanan grid. We round the y-coordinate of every segment endpoint in A' down to the next multiple of N^4 (possibly removing segments of length 0). The resulting drawing remains connected; every path to the origin remains monotone; and since the segments of A' lie on the Hanan grid of $S' \cup \{(0, 0)\}$, no new cycles are introduced. Thus, the resulting drawing constitutes an arborescence A'' for the set S'' of sinks obtained by scaling S by N^4. Since A' lies on the Hanan grid, it is a union of N paths, each with at most N vertical segments. The rounding operation increases the length of each such vertical segment by at most N. Thus, the total length of A'' is at most $kN^4 + 2N^3$. By Theorem 2.1 there exists an optimum arborescence A^* for S'' that lies on the Hanan grid. The length of A^* is a multiple of N^4, and thus at most kN^4, since $2N^3 < N^4$ for $N \geq 3$. It follows that S has an RSA of length at most k.

Therefore, (S, k) is a yes-instance for Rectilinear Steiner Arborescence if and only if $(S', kN^4 + N^3)$ is a yes-instance for YRSA. Since $(S', kN^4 + N^3)$ can be computed in polynomial time from (S, k), and since the coordinates in S' are polynomially bounded in the coordinates of S, it follows that YRSA is strongly NP-complete. □

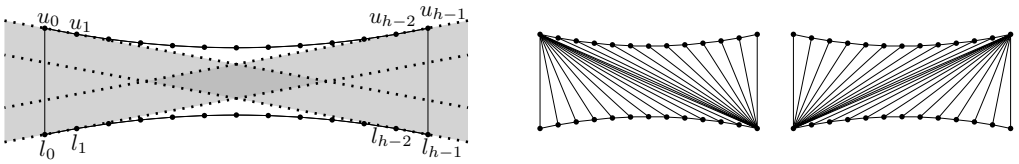

Fig. 2 Left: The polygon and the hourglass (gray) of a double chain. The diamond-shaped flip-kernel can be extended arbitrarily by flattening the chains. Right: The upper extreme triangulation T_u and the lower extreme triangulation T_l.

Due to Theorem 2.1, we get the following technical corollary, which will be useful later.

Corollary 2.3 *YRSA remains strongly NP-complete even if the sinks have coordinates that are a multiple of a positive integer whose value is polynomial in N.*

3 Double Chains

Our definitions (and illustrations) follow [19]. A *double chain* D is a polygon that consists of two chains, an *upper chain* and a *lower chain*. There are h vertices on each chain, $\langle u_0, \ldots, u_{h-1}\rangle$ on the upper chain and $\langle l_0, \ldots, l_{h-1}\rangle$ on the lower chain, both numbered from left to right, and D is defined by $\langle l_0, \ldots, l_{h-1}, u_{h-1}, \ldots, u_0\rangle$. Any point on one chain sees every point on the other chain, and any quadrilateral formed by three vertices of one chain and one vertex of the other chain is non-convex; see Fig. 2 (left). We call the triangulation T_u of D where u_0 has maximum degree the *upper extreme triangulation*; observe that this triangulation is unique. The triangulation T_l of D where l_0 has maximum degree is called the *lower extreme triangulation*. The two extreme triangulations are used to show that the diameter of the flip graph is quadratic; see Fig. 2 (right).

Theorem 3.1 (Hurtado, Noy, Urrutia [11]) *The flip distance between T_u and T_l is $(h-1)^2$.* □

Through a slight modification of D, we can make the flip distance between the upper and the lower extreme triangulation linear. This will enable us in our reduction to impose a certain structure on short flip sequences. To describe this modification, we first define the *flip-kernel* of a double chain.

Let W_1 be the wedge defined by the lines through u_0u_1 and l_0l_1 whose interior contains no vertex of D but intersects the segment u_0l_0. Define W_h analogously by the

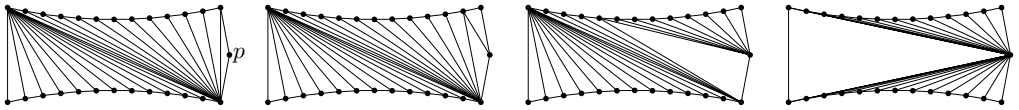

Fig. 3 The extra point p in the flip-kernel of D allows flipping one extreme triangulation of Q to the other in $4h-4$ flips.

lines through $u_{h-1}u_{h-2}$ and $l_{h-1}l_{h-2}$. We call $W := W_1 \cup W_h$ the *hourglass of* D. The unbounded set $W \cup D$ is defined by four rays and the two chains. The *flip-kernel* of D is the intersection of the four closed half-planes below the lines through u_0u_1 and $u_{h-2}u_{h-1}$ and above the lines through l_0l_1 and $l_{h-2}l_{h-1}$.[2]

Definition 3.2 Let D be a double chain and let p be a point in the flip-kernel of D to the right of the directed line $l_{h-1}u_{h-1}$. The polygon given by the sequence $\langle l_0, \ldots, l_{h-1}, p, u_{h-1}, \ldots, u_0\rangle$ is called a *double chain extended by* p. The upper and the lower extreme triangulation of such a polygon contain the edge $u_{h-1}l_{h-1}$ as a diagonal and are otherwise defined in the same way as for D.

The flip distance between the two extreme triangulations of D extended by a point p is much smaller than for D [24]. Fig. 3 shows how to transform them into each other with $4h-4$ flips. The next lemma shows that this is optimal, even for more general polygons. The lemma is a slight generalization of a lemma by Lubiw and Pathak [18] on double chains of constant size.

Lemma 3.3 *Suppose that $h \geq 5$ and consider a polygon that contains D and has $\langle l_0, \ldots, l_{h-1}\rangle$ and $\langle u_{h-1}, \ldots, u_0\rangle$ as part of its boundary. Let T_1 and T_2 be two triangulations that contain the upper extreme triangulation and the lower extreme triangulation of D as a sub-triangulation, respectively. Then T_1 and T_2 have flip distance at least $4h-4$.*

Proof We slightly generalize a proof by Lubiw and Pathak [18] for double chains of constant size.

Let C_u be the upper chain and C_l be the lower chain of D. The triangulation T_1 has $2(h-1)$ triangles with an edge on C_u or on C_l. These triangles are called *anchored*, and the vertex not incident to the edge on C_u or on C_l is called the *apex*. For each anchored triangle with an edge on C_u, the apex must move from l_{h-1} to l_0, and similarly for C_l.

[2] The flip-kernel of D might not be completely inside the polygon D. This is in contrast to the "visibility kernel" of a polygon.

We distinguish three types of flips depending on whether the convex quadrilateral whose diagonal is flipped has (1) four; (2) three; or (3) at most two vertices on D. A flip of type (1) moves the apex of two anchored triangles by one; a flip of type (2) moves the apex of one anchored triangle from D to a point outside D or back again; and a flip of type (3) does not move any apex of an anchored triangle along D or between a vertex of D and a vertex not in D.

We say that an anchored triangle is of type (1) if its apex is moved only by flips of type (1). It is of type (2) if its apex is moved by at least one flip of type (2). Every anchored triangle is either of type (1) or of type (2). A type (1) triangle must be involved in at least $h-1$ flips of type (1), and each of these flips can affect at most one other type (1) triangle. A type (2) triangle must be involved in at least 2 flips of type (2), and each of these flips can affect no other anchored triangle. Thus, if we have m_1 type (1) triangles and m_2 type (2) triangles, we need at least $(h-1)m_1/2+2m_2$ flips. For $h \geq 5$, we have $(h-1)m_1/2 + 2m_2 \geq 2(m_1+m_2) = 4h-4$, as claimed. □

The following result can be seen as a special case of [19, Proposition 1].

Lemma 3.4 *Consider a polygon that contains D and has $\langle u_{h-1}, \ldots, u_0, l_0, \ldots, l_{h-1}\rangle$ as part of its boundary. Let T_1 and T_2 be two triangulations that contain the upper and the lower extreme triangulation of D as a sub-triangulation, respectively. Let σ be a flip sequence from T_1 to T_2 such that there is no triangulation in σ containing a triangle with one vertex at the upper chain, the other vertex at the lower chain, and the third vertex at a point in the interior of the hourglass of D. Then $|\sigma| \geq (h-1)^2$.*

Proof Our reasoning is similar to the proof of Lemma 3.3, see also [18]. As before, let C_u and C_l be the upper and lower chain of D, and call a triangle with an edge on C_u or on C_l *anchored*, the third vertex being the *apex*. Any triangulation of the given polygon has $2(h-1)$ anchored triangles.

We will argue that for each triangulation of σ there exists a line ℓ that separates C_u from C_l and that intersects all anchored triangles. This is clear if the apices of all anchored triangles lie on the other chain or outside the hourglass. Now consider a triangulation of the sequence σ where at least one anchored triangle has its apex at a vertex v inside the hourglass. Let r be a ray that starts at a point on u_0l_0 and passes through v such that the supporting line of r separates C_u from C_l (such a ray must exist since v is inside the hourglass). Then r intersects at least one triangle that is not

anchored, because the triangle whose interior is intersected by r before reaching v cannot be anchored. Let Δ be the first non-anchored triangle whose interior is intersected by r. Then Δ has one vertex on C_u and one vertex on C_l. By assumption, the third vertex of Δ cannot be inside the hourglass, so it must lie outside. This means that one of the vertices of Δ has to be either u_{h-1} or l_{h-1}. This implies that either all anchored triangles at C_u or or all anchored triangles C_l, respectively, have their apex at the opposite chain. Thus, also for this triangulation there exists a line ℓ that separates C_u from C_l and that intersects all anchored triangles. Observe that every such line intersects the anchored triangles in the same order.

Now we proceed similarly as in the proof of Hurtado, Noy, and Urrutia [11]: we observe that an anchored triangle at C_u and an anchored triangle at C_l can change their relative position along ℓ only if they have an edge in common and this edge is flipped. This results in an overall number of $(h-1)^2$ flips. □

4 The Reduction

We reduce YRSA to PolyFlip. Let S be a set of N sinks. By Corollary 2.3, we can assume that the coordinates of sinks of S are multiples of a factor $\beta = 2N$ in $\{0, \ldots, \beta n\}$. Further, we can restrict ourselves to YRSA instances of the form $(S, \beta k)$. Thus, we imagine that the sinks are embedded on a $\beta n \times \beta n$ grid. The reasons for the choice of β will become clear below.

We construct a polygon P and two triangulations T_1, T_2 in P such that a shortest flip sequence from T_1 to T_2 corresponds to a shortest RSA for S. To this end, we will describe how to interpret any triangulation of P as a *chain path*, a path in the integer grid that starts at the root and uses only edges that go north or east. It will turn out that flips in P essentially correspond to moving the endpoint of the chain path along the grid. We choose P, T_1, and T_2 in such a way that a shortest flip sequence between T_1 and T_2 moves the endpoint of the chain path according to an Eulerian traversal of a shortest RSA for S. To force the chain path to visit all sinks, we use the observations from Section 3: the polygon P contains a double chain for each sink, so that only for certain triangulations of P it is possible to flip the double chain quickly. These triangulations will be exactly the triangulations that correspond to the chain path visiting the appropriate sink. To force the sinks to be visited, we, with foresight, fix the number of points in

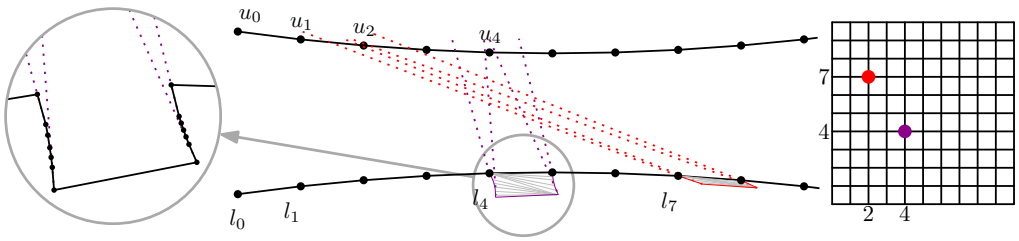

Fig. 4 The sink gadget for a sink (x_s, y_s) is obtained by replacing the edge $l_{y_s} l_{y_s+1}$ by a double chain with d vertices on each chain. The double chain is oriented such that u_{x_s} is the only point inside its hourglass and its flip-kernel.

each of the two chains of a double chain representing a sink to $d = nN$ (recall that n is polynomial in N).

4.1 The Construction

We take a double chain D with $\beta n + 2$ vertices on each chain such that the flip-kernel of D extends to the right of $l_{\beta n+1} u_{\beta n+1}$. We add a point z to that part of the flip-kernel, and we let Q be the polygon defined by $\langle l_0, \ldots, l_{\beta n+1}, z, u_{\beta n+1}, \ldots, u_0 \rangle$, i.e., a double chain extended by z (recall Definition 3.2). Next, we add double chains to Q in order to encode the sinks in S. For each sink $s = (x_s, y_s)$, we remove the edge $l_{y_s} l_{y_s+1}$, and we replace it by a (rotated) double chain D_s with d vertices on each chain, such that l_{y_s} and l_{y_s+1} become the last point on the lower and the upper chain of D_s, respectively. We orient D_s in such a way that u_{x_s} is the only point inside the hourglass of D_s and so that u_{x_s} lies in the flip-kernel of D_s; see Fig. 4. We refer to the added double chains as the *sink gadgets*, and we call the resulting polygon P. Since the y-coordinates in S are pairwise distinct, there is at most one sink gadget per edge of the lower chain of Q. Since $\beta \geq 2$, no two sink gadgets are placed on neighboring edges of Q, and can be constructed such that they do not overlap. Hence, P is a simple polygon. The precise placement of the sink gadgets is flexible, so given an appropriate embedding of D, we can make all coordinates integers whose value is polynomial in the input size; see Appendix A for details.

Next, we describe the source and target triangulation for P. In the source triangulation T_1, the interior of Q is triangulated such that all edges are incident to z. The sink gadgets are all triangulated with the upper extreme triangulation. The target triangula-

tion T_2 is similar, but now the sink gadgets are all triangulated with the lower extreme triangulation.

To get from T_1 to T_2, we must go from one extreme triangulation to the other for each sink gadget D_s. By Lemma 3.4, this requires $(d-1)^2$ flips, unless the flip sequence creates a triangle that allows us to use the vertex in the flip-kernel of D_s. In this case, we say that the flip sequence *visits* the sink s. The main idea is that, since the value chosen for d is large, a shortest flip sequence must visit all sinks, and we will show that this induces an RSA for S of comparable length. Conversely, we will show how to derive a flip sequence from an RSA. The precise statement is given in the following theorem.

Theorem 4.1 *Let $N \geq 3$, and set $\beta = 2N$. Let S be a set of N sinks such that the coordinates of the sinks are multiples of β in $\{0, \ldots, \beta n\}$, where n is polynomially bounded in N. Set $d = nN$ and let P be the simple polygon and T_1 and T_2 the two triangulations of P as described above. Then for any $k \geq 1$, the flip distance between T_1 and T_2 w.r.t. P is at most $2\beta k + (4d-2)N$ if and only if S has an RSA of length at most βk.*

We will prove Theorem 4.1 in the following sections. But first, let us show how to use it for our NP-completeness result.

Theorem 4.2 PolyFlip *is* NP*-complete.*

Proof As mentioned in the introduction, the flip distance in polygons is polynomially bounded, so PolyFlip is in NP. We reduce from YRSA. Let $(S, \beta k)$ be an instance of YRSA as above. We construct P and T_1, T_2 as described above. This takes polynomial time (see Appendix A for details on the coordinate representation). By Theorem 4.1, there exists an RSA for S of length at most βk if and only if there exists a flip sequence between T_1 and T_2 of length at most $2\beta k + (4d-2)N$. □

4.2 Chain Paths

Now we introduce the *chain path*, our main tool to establish a correspondence between flip sequences and RSAs. Let T be a triangulation of Q (i.e., the polygon P without the sink gadgets, cf. Section 4.1). A *chain edge* is an edge of T between the upper and the lower chain of Q. A *chain triangle* is a triangle of T that contains two chain edges. Let $e_1, \ldots, e_m$ be the chain edges, sorted from left to right according to their intersection

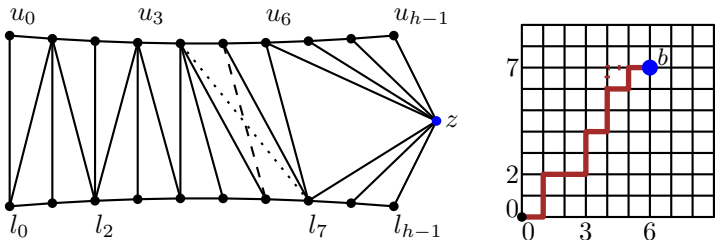

Fig. 5 A triangulation of Q and its chain path. Flipping edges to and from z moves the endpoint b along the grid. A flip between chain triangles changes a bend.

with a line that separates the upper from the lower chain. For $i = 1, \ldots, m$, write $e_i = (u_v, l_w)$ and set $c_i = (v, w)$. In particular, $c_1 = (0, 0)$. Since T is a triangulation, any two consecutive edges e_i, e_{i+1} share one endpoint, while the other endpoints are adjacent on the corresponding chain. Thus, c_{i+1} dominates c_i and $\|c_{i+1} - c_i\|_1 = 1$. It follows that $c_1 c_2 \ldots c_m$ is an x- and y-monotone path, beginning at the root. It is called the *chain path* for T. Each vertex of the chain path corresponds to a chain edge, and each edge of the chain path corresponds to a chain triangle. Conversely, every chain path induces a triangulation T of Q; see Fig. 5. In the following, we let b denote the upper right endpoint of the chain path. We now investigate how flipping edges in T affects the chain path.

Observation 4.3 *Suppose we flip an edge that is incident to z. Then the chain path is extended by moving b north or east.* □

Observation 4.4 *Suppose that T contains at least one chain triangle. When we flip the rightmost chain edge, we shorten the chain path at b.* □

Finally, we can flip an edge between two chain triangles. This operation is called a *chain flip*.

Observation 4.5 *A chain flip changes a bend from east to north to a bend from north to east, or vice versa.*

Proof If a chain edge $u_i l_j$ is incident to two chain triangles and is flippable, then the two triangles must be of the form $u_i u_{i-1} l_j$ and $l_j l_{j+1} u_i$, or $u_{i+1} u_i l_j$ and $l_{j-1} l_j u_i$. Thus, flipping $u_i l_j$ corresponds exactly to the claimed change in the chain path. □

Corollary 4.6 *A chain flip does not change the length of the chain path.* □

We summarize the results of this section in the following lemma:

Lemma 4.7 *Any triangulation T of Q uniquely determines a chain path, and vice versa. A flip in T corresponds to one of the following operations on the chain path: (i) move the endpoint b north or east; (ii) shorten the path at b; (iii) change an east-north bend to a north-east bend, or vice versa.* □

4.3 From an RSA to a Short Flip Sequence

Using the notion of a chain path, we now prove the "if" direction of Theorem 4.1.

Lemma 4.8 *Let $k \geq 1$ and A an RSA for S of length βk. Then the flip distance between T_1 and T_2 w.r.t. P is at most $2\beta k + (4d-2)N$.*

Proof The triangulations T_1 and T_2 both contain a triangulation of Q whose chain path has its endpoint b at the root. We use Lemma 4.7 to generate flips inside Q so that b traverses A in a depth-first manner. This needs $2\beta k$ flips.

Each time b reaches a sink s, we move b north. This creates a chain triangle that allows the edges in the sink gadget D_s to be flipped to the auxiliary vertex in the flip-kernel of D_s. The triangulation of D_s can then be changed with $4d-4$ flips; see Lemma 3.3. Next, we move b back south and continue the traversal. Moving b at s needs two additional flips, so we take $4d-2$ flips per sink, for a total of $2\beta k + (4d-2)N$ flips. □

4.4 From a Short Flip Sequence to an RSA

Finally, we consider the "only if" direction in Theorem 4.1. Let τ be a flip sequence on Q. We say that τ *visits* a sink $s = (x_s, y_s)$ if τ has at least one triangulation that contains the chain triangle $u_{x_s} l_{y_s} l_{y_s+1}$. We call τ a *flip traversal* for S if (i) τ begins and ends in the triangulation whose corresponding chain path has its endpoint b at the root and (ii) τ visits every sink in S. The following lemma shows that every short flip sequence σ in P can be mapped to a flip traversal (where with "short", we mean $|\sigma| < (d-1)^2$).

Lemma 4.9 *Let σ be a flip sequence from T_1 to T_2 w.r.t. P with $|\sigma| < (d-1)^2$. Then there is a flip traversal τ for S with $|\tau| \leq |\sigma| - (4d-4)N$.*

Proof We show how to obtain a flip traversal τ for S from σ. Let T be a triangulation of P. A triangle of T is an *inner triangle* if all its sides are diagonals. It is an *ear* if two

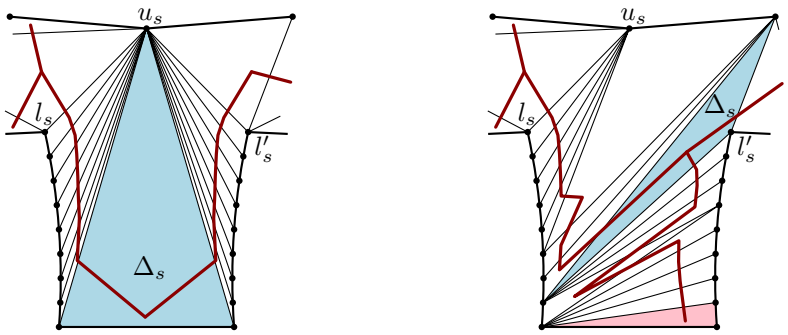

Fig. 6 Triangulations of D_s in P with $\Delta_s = \Delta$ (left), and with Δ being an ear (red) and Δ_s an inner triangle (right). The fat tree indicates the dual.

of its sides are polygon edges. By construction, every inner triangle of T must have (i) one vertex incident to z (the rightmost vertex of Q), or (ii) two vertices incident to a sink gadget (or both). There can be only one triangle of type (ii) per sink gadget. The weak (graph theoretic) dual of T is a tree in which ears correspond to leaves and inner triangles have degree 3.

For a sink $s = (x_s, y_s)$, let D_s be the corresponding sink gadget. It lies between the vertices l_{y_s} and l_{y_s+1} and has exactly u_{x_s} in its flip kernel. For brevity, we will write l_s for l_{y_s}, l'_s for l_{y_s+1}, and u_s for u_{x_s}. We define a triangle Δ_s for D_s. Consider the bottommost edge e of D_s, and let Δ be the triangle of T that is incident to e. By construction, Δ is either an ear of T, or it is the triangle defined by e and u_s. In the latter case, we set $\Delta_s = \Delta$. In the former case, we claim that T has an inner triangle Δ' with two vertices on D_s: follow the path from Δ in the weak dual of T; while the path does not encounter an inner triangle, the next triangle must have an edge of D_s as a side. There is only a limited number of such edges, so eventually we must meet an inner triangle Δ'. We then set $\Delta_s = \Delta'$; see Fig. 6. Note that Δ_s might be $l_s l'_s u_s$.

For each sink s, let the polygon Q_s consist of D_s extended by the vertex u_s (cf. Definition 3.2). Let T be a triangulation of P. We show how to map T to a triangulation T_Q of Q and to triangulations T_s of Q_s, for each s.

We first describe T_Q. It contains every triangle of T with all three vertices in Q. For each triangle Δ in T with two vertices on Q and one vertex on the left chain of a sink gadget D_s, we replace the vertex on D_s by l_s. Similarly, if the third vertex of Δ is on the right chain of D_s, we replace it by l'_s. For every sink s, the triangle Δ_s has one vertex

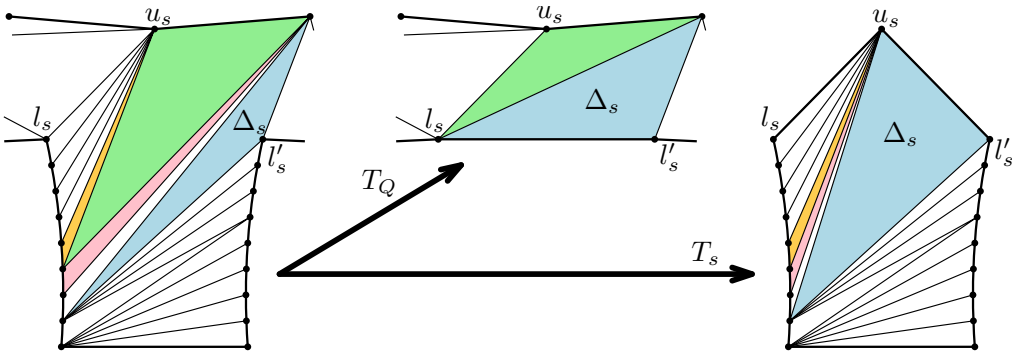

Fig. 7 Obtaining T_Q and T_s from T.

at a point u_i of the upper chain. In T_Q, we replace Δ_s by the triangle $l_s l'_s u_i$. No two triangles in T_Q overlap, and they cover all of Q. Thus, T_Q is indeed a triangulation of Q.

Now we describe how to obtain T_s, for a sink $s \in S$. Each triangle of T with all vertices on Q_s is also in T_s. Each triangle with two vertices on D_s and one vertex not in Q_s is replaced in T_s by a triangle whose third vertex is moved to u_s in T_s (note that this includes Δ_s); see Fig. 7. Again, all triangles cover Q_s and no two triangles overlap.

Finally, we show that a flip in T corresponds to at most one flip either in T_Q or in precisely one T_s for some sink s. We do this by considering all the possibilities for two triangles that share a common flippable edge. By construction, no two triangles that are mapped to two different triangulations T_s and T_t for sinks $s \neq t \in S$ can share an edge.

Case 1. We flip an edge between two triangles that are either both mapped to T_Q or to T_s and are different from Δ_s. This flip clearly happens in at most one triangulation.

Case 2. We flip an edge between a triangle Δ_1 that is mapped to T_s and a triangle Δ_2 that is mapped to T_Q, such that both Δ_1 and Δ_2 are different from Δ_s. This results in a triangle Δ'_1 that is incident to the same edge of Q_s as Δ_1, and a triangle Δ'_2 having the same vertices of Q as Δ_2. Since the apex of Δ_1 is a vertex of the upper chain or z (otherwise, it would not share an edge with Δ_2), it is mapped to u_s, as is the apex of Δ'_1. Also, the apex of Δ'_2 is on the same chain of D_s as the one of Δ_2. Hence, the flip affects neither T_Q nor T_s.

Case 3. We flip the edge between a triangle Δ_2 mapped to T_Q and Δ_s. By construction, this can only happen if Δ_s is an inner triangle. The flip affects only T_Q, because the new inner triangle Δ'_s is mapped to the same triangle in T_s as Δ_s, since both apexes are moved to u_s.

Case 4. We flip the edge between a triangle Δ of T_s and Δ_s. Similar to Case 3, this affects only T_s, because the new triangle Δ'_s is mapped to the same triangle in T_Q as Δ_s, since the two corners are always mapped to l_s and l'_s.

Thus, σ induces a flip sequence τ in Q and flip sequences σ_s in each Q_s so that $|\tau| + \sum_{s\in S} |\sigma_s| \leq |\sigma|$. Furthermore, each flip sequence σ_s transforms Q_s from one extreme triangulation to the other. Since $|\sigma| < (d-1)^2$, Lemma 3.4 tells us that the triangulations T_s have to be transformed so that Δ_s has a vertex at u_s at some point. Moreover, by Lemma 3.3, we have $|\sigma_s| \geq 4d-4$ for each $s \in S$. Thus, τ is a flip traversal, and $|\tau| \leq |\sigma| - N(4d-4)$, as claimed. □

In order to obtain a static RSA from a changing flip traversal, we use the notion of a *trace*. A *trace* is a domain on the grid. It consists of *edges* and *boxes*: an edge is a line segment of length 1 whose endpoints have positive integer coordinates; a box is a square of side length 1 whose corners have positive integer coordinates. Similar to arborescences, we require that a trace R (i) is (topologically) connected; (ii) contains the root $(0,0)$; and (iii) from every grid point contained in R there exists an x- and y-monotone path to the root that lies completely in R. We say R is a *covering trace* for S (or, R *covers* S) if every sink in S is covered by R (i.e., incident to a box or an edge in R).

Let τ be a flip traversal as in Lemma 4.9. By Lemma 4.7, each triangulation in τ corresponds to a chain path. This gives a covering trace R for S in the following way. For every flip in τ that extends the chain path, we add the corresponding edge to R. For every flip in τ that changes a bend, we add the corresponding box to R. Afterwards, we remove from R all edges that coincide with a side of a box in R. Clearly, R is (topologically) connected. Since τ is a flip traversal for S, every sink is covered by R. Note that every grid point p in R is connected to the root by an x- and y-monotone path on R, since at some point p belonged to a chain path in τ. Hence, R is indeed a trace, the unique *trace of* τ. Note that not only a flip traversal but any flip sequence starting with a zero-length chain path defines a trace in this way.

Next, we define the *cost* of a trace R, $\mathrm{cost}(R)$, so that if R is the trace of a flip traversal τ, then $\mathrm{cost}(R)$ gives a lower bound on $|\tau|$. An edge has cost 2. Let B be a box in R. A *boundary side* of B is a side that is not part of another box. The cost of B is 1 plus the number of boundary sides of B. Then, $\mathrm{cost}(R)$ is the total cost over all boxes

and edges in R. For example, the cost of a tree is twice the number of its edges, and the cost of a rectangle is its area plus its perimeter. An edge can be interpreted as a degenerated box, having two boundary sides and no interior.

Proposition 4.10 *Let τ be a flip traversal and R the trace of τ. Then cost$(R) \leq |\tau|$.*

Proof Let ς_i be the sequence of the first i triangulations of τ, R_i the trace defined by ς_i, and let κ_i be the length of the chain path for the ith triangulation. We will show by induction on i that $\text{cost}(R_i) \leq i + \kappa_i$, for $i = 1, \ldots, |\tau|$. Since $\varsigma_{|\tau|} = \tau$, $R_{|\tau|} = R$, and $\kappa_{|\tau|} = 0$, this gives the desired result.

After the first flip, R_1 is an edge (so $\text{cost}(R_1) = 2$), and $\kappa_1 = 1$, which fulfills the invariant. Consider the ith flip. If the flip extends the chain path, the cost of the trace increases by at most 2, and the length of the chain path increases by 1, fulfilling the invariant. If the flip contracts the chain path, the trace does not change, but the length of the chain path is decreased by 1, again fulfilling the invariant. We are therefore left with the case where the flip is a chain flip. We have $\kappa_{i-1} = \kappa_i$, so we have to show that $\text{cost}(R_i) \leq \text{cost}(R_{i-1}) + 1$. We may assume that the flip adds a box B to R_{i-1} (otherwise, the cost of the trace remains unchanged). Consider the intersection of the boundary of B with the one of R_{i-1}. This intersection contains at least two elements, as the chain path is part of R_{i-1}. An edge in the intersection becomes a boundary side in R_i, reducing the cost by 1. A boundary side in the intersection vanishes in R_i, also reducing the cost by 1. Thus, adding B creates a box and at most two boundary sides, causing a cost of at most 3, but it simultaneously reduces the cost by at least 2. See the examples in Fig. 8. The overall cost increases at most by 1, and the invariant is maintained. □

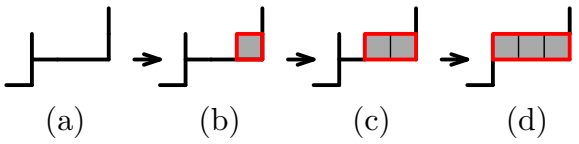

Fig. 8 Examples of how boundary sides (red) are added to a trace. To a trace of cost 16 (a) a box (gray) is added (b), which transforms two edges in boundary sides and adds two boundary sides, resulting in an overall cost of 17. The next box removes one boundary side and one edge and adds three boundary sides (c), the cost becomes 18. A box might also remove more than two elements (d), reducing the overall cost to 17.

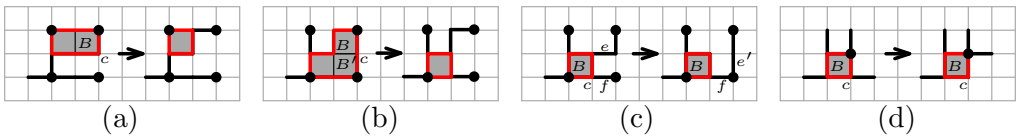

Fig. 9 Parts of traces to be modified; the boundary sides are shown in red. (a) A box that has a corner c with no incident elements can be removed. (b) Two adjacent boxes that have a shared corner c without any incident elements can be removed. (c) Replacing a single edge. (d) Sliding an edge.

Now we relate the length of an RSA for S to the cost of a covering trace for S, and thus to the length of a flip traversal. Since each sink is connected in R to the root by an x- and y-monotone path, traces can be regarded as generalized RSAs. In particular, we make the following observation.

Observation 4.11 *Let R be a covering trace for S that contains no boxes, and let A_τ be a shortest path tree in R from the root to all sinks in S. Then A_τ is an RSA for S.* □

If τ contains no flips that change bends, the corresponding trace R has no boxes. Then, R contains an RSA A_τ with $2|A_\tau| \leq \text{cost}(R)$, by Observation 4.11. The next lemma shows that, due to the fact that β is even, there is always a shortest covering trace for S that does not contain any boxes.

Lemma 4.12 *Let τ be a flip traversal of S. Then there exists a covering trace R for S such that R does not contain a box and such that cost$(R) \leq |\tau|$.*

To prove the lemma, we investigate the structure of minimal covering traces. There exists at least one trace of cost at most $|\tau|$, namely the trace of τ. Let $\mathcal{R}_1$ be the set of all covering traces for S that have minimum cost. Let $\mathcal{R}_2 \subseteq \mathcal{R}_1$ be those covering traces among $\mathcal{R}_1$ that contain the minimum number of boxes. If $\mathcal{R}_2$ contains a trace without boxes, we are done, as every covering trace in $\mathcal{R}_2$ fulfills the requirements of Lemma 4.12. We show that this is actually the case by assuming, for the sake of contradiction, that every covering trace in $\mathcal{R}_2$ contains at least one box.

Let $R \in \mathcal{R}_2$ and suppose that R contains a box. Let B be a *maximal* box in R, i.e., R has no other box whose lower left corner has both x- and y-coordinate at least as large as the lower left corner of B. In order to prove Lemma 4.12, we need several lemmata on traces of minimum cost.

Lemma 4.13 *Let B be a maximal box and let c a corner of B that is not the root $(0, 0)$. Then c is incident either to a sink, an edge, or another box.*

Proof Suppose there exists a corner c for which this is not the case. Note that such a c cannot be the lower left corner of B, as there has to be an x- and y-monotone path to the root. Hence, we could remove c and B while keeping the sides of B not incident to c as edges, if necessary; see Fig. 9(a). In the resulting structure, every element still has an x- and y-monotone path to the root: If c is the lower right or upper left corner, any path initially passing through c could be rerouted to pass through the corner opposite of c in B. If c is the upper right corner of B, no path is passing through c. Hence, the resulting structure would be a covering trace with smaller cost, contradicting the choice of R. □

Lemma 4.14 *Suppose B shares a horizontal side with another box B'. Let c be the right endpoint of the common side. Then c is incident either to a sink, an edge, or another box.*

Proof Suppose this is not the case. Then we could remove B and B' from R while keeping the sides not incident to c as edges, if necessary; see Fig. 9(b). This results in a valid trace that has no higher cost but less boxes than R, contradicting the choice of R. □

Lemma 4.15 *Let c be the lower right corner of B. Then c has no incident vertical edge.*

Proof Such an edge would be redundant, since c already has an x- and y-monotone path to the root that goes through the lower left corner of B. □

Proof (Proof of Lemma 4.12) Using the Lemmata 4.13, 4.14, and 4.15, we derive a contradiction from the choice of R and the maximal box B. Note that since β is even, all sinks in S have even x- and y-coordinates. We distinguish two cases.

Case 1. There exists a maximal box B whose top right corner c' does not have both coordinates even. Suppose that the x-coordinate of c' is odd. (Otherwise, mirror the plane at the line $x = y$ to swap the x- and the y-axis. Note that the property of being a trace is invariant under mirroring the plane along the line $x = y$; in particular, the choice of B in R as a maximal box remains valid) By Lemma 4.13, there is at least one edge incident to the top right corner of B (it cannot be a box by the choice of B, and it cannot be a sink because of the current case). Recall the slide operation for an edge in an arborescence. This operation can easily be adapted in an analogous way to traces. If there is a vertical edge v incident to c', it cannot be incident to a sink. Thus,

we could slide v to the right (together with all other vertical edges that are above v and on the supporting line of v). Hence, we may assume that c' is incident to a single horizontal edge e; see Fig. 9(c). By Lemma 4.13, the bottom right corner c of B must be incident to an element. We know that c cannot be the top right corner of another box (Lemma 4.14), nor can it be incident to a vertical segment (Lemma 4.15). Thus, c is incident to an element f that is either a horizontal edge or a box with top left corner c. But then e could be replaced by a vertical segment e' incident to f, and afterwards B could be removed as in the proof of Lemma 4.13, contradicting the choice of R.

Case 2. The top right corner of each maximal box has even coordinates. Let B be the rightmost maximal box. As before, let c be the bottom right corner of B. The y-coordinate of c is odd; see Fig. 9(d). By the choice of B, we know that c is not the top left corner of another box: this would imply that there is another maximal box to the right of B. We may assume that c is not incident to a horizontal edge, as we could slide such an edge up, as in Case 1. Furthermore, c cannot be incident to a vertical edge (Lemma 4.15), nor be the top right corner of another box (Lemma 4.14). Thus, B violates Lemma 4.13, and Case 2 also leads to a contradiction.

Thus, the choice of R forces a contradiction in either case. Hence, the minimum number of boxes in a minimum covering trace for S is 0. □

Now we can finally complete the proof of Theorem 4.1 by giving the second direction of the correspondence.

Lemma 4.16 *Let $k \geq 1$ and let σ be a flip sequence on P from T_1 to T_2 with $|\sigma| \leq 2\beta k + (4d-2)N$. Then there exists an RSA for S of length at most βk.*

Proof Trivially, there always exists an RSA on S of length less than $2\beta nN$, so we may assume that $k < 2nN$. Hence (recall that $\beta = 2N$ and $d = nN$),

$$2\beta k + (4d-2)N < 2 \cdot 2N \cdot 2nN + 4nN^2 - 2N < 12nN^2 < (d-1)^2,$$

for $n \geq 14$ and positive N. Thus, σ meets the requirements of Lemma 4.9, and therefore we can obtain a flip traversal τ for S with $|\tau| \leq 2\beta k + 2N$. By Lemma 4.12 and Observation 4.11, we can conclude that there is an RSA A for S that has length at most $\beta k + N$. By Theorem 2.1, there is an RSA A' for S that is not longer than A and that lies on the Hanan grid for S. The length of A' must be a multiple of β. Thus, since $\beta > N$, we get that A' has length at most βk. □

5 Conclusion

In this paper, we showed NP-hardness of determining a shortest flip sequence between two triangulations of a simple polygon. This complements the recent hardness results for point sets (obtained by reduction from variants of Vertex Cover). However, while for point sets the problem is hard to approximate as well, our reduction does not rule out the existence of a polynomial-time approximation scheme (PTAS), since a PTAS is known for the RSA problem [17]. When problems that are hard for point sets are restricted to simple polygons, the application of standard techniques—like dynamic programming—often gives polynomial-time algorithms. This is, for example, the case for the construction of the minimum weight triangulation. Our result illustrates that determining the flip distance is a different, harder type of problem. Is there a PTAS for the flip distance between triangulations of a polygon? Even a constant-factor approximation would be interesting.

For convex polygons (or, equivalently, points in convex position), the complexity of the problem remains unknown. Our construction heavily relies on the double chain construction, using many reflex vertices. Does the problem remain hard if we restrict the number of reflex vertices to some constant fraction?

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A A Note on Coordinate Representation

Since it is necessary for the validity of the proof that the input polygon can be represented in size polynomial in the size of the YRSA instance, we give a possible method to embed the polygon with vertices at rational coordinates whose numerator and denominator are polynomial in N. By an additional perturbation argument we can guarantee integer coordinates whose values are polynomial in N (which slightly strengthens the result). We first introduce the general technique used for the embedding, and then give further details on how the sink gadgets are constructed (using methods similar to [19]). Finally, we explain how the construction can be transformed to integer points in general position.

A.1 Placing Points on Arcs

The main idea of the construction is to place all vertices on rational points on circular arcs. There are two large arcs where we place the vertices of the upper and the lower chain, and smaller arcs on which we place the vertices of the sink gadgets. All these circular arcs are chosen from *rational circles*, i.e., circles that are defined by three rational points. Similarly, a *rational line* is a line trough a rational point with rational slope (or, equivalently, a line defined by two rational points). It is well-known that, if one of the two intersection points of a rational line with a rational circle is a rational point, then the other intersection point is rational as well (see, e.g., [12, p. 5]). Hence, given a rational point p on a rational circle, we can obtain an arbitrary number of rational points on the circle via different rational lines through p.

Let us apply this for one possible way of constructing the double chain D. The construction is shown in Fig. 10 (left). We place the $\beta n + 2$ points of the lower chain on the unit circle (with center at the origin). Let ℓ_i be the line through $(-1, 0)$ with slope $1 + \frac{i}{\beta n+2}$. For $i = 1, \ldots, \beta n/2 + 1$, we get $\beta n/2 + 1$ rational points on the upper-left quadrant of the unit circle from the intersections with this family of lines. We can do the analogous construction for points on the upper-right quadrant by choosing lines through $(1, 0)$ with a negative slope $(-1) - \frac{i}{\beta n+2}$. In this way, we obtain the vertices of the lower chain of D. For the upper chain, we place points on the unit circle with origin $(0, 3)$ analogously. Note that line $\ell_{\beta n/2+1}$ passes through $(1, 3)$, so when picking rational points on the lower-right and lower-left quadrant of the second unit circle for the upper chain, the resulting point set is indeed the vertex set of a double chain in which the line through l_0 and $u_{\beta n+1}$ is $\ell_{\beta n/2+1}$. Finally, note that all slopes used in the construction have numerators and denominators that are polynomial in N. Hence, this also holds for the coordinates of the vertices of D. Note that this is, essentially, the parametrization of the unit circle, as discussed in [5].

Clearly, this method is not restricted to unit circles. We now discuss the following main building block for constructing the sink gadgets. Given three rational points p_1, p_2, p_3, we construct a circular arc on a rational circle that starts at p_1, ends at p_2 and is completely contained inside the triangle $p_1p_2p_3$. Then, we choose an arbitrary number of rational points on that circular arc. This is illustrated in Fig. 10 (right). W.l.o.g, let the inner angle of the triangle at p_1 be less than or equal to the one at p_2. Let Z be the circle through p_1 and p_2 such that the line p_1p_3 is a tangent of Z. Clearly, Z is well-defined, and the arc between p_1 and p_2 is inside the triangle. The circle Z is rational. (Consider

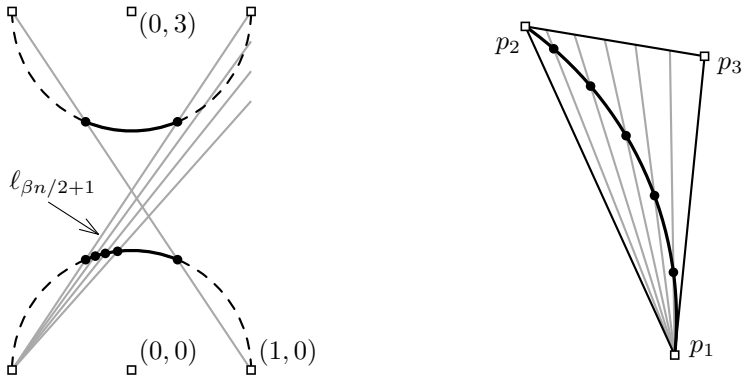

Fig. 10 Left: Construction of the main double chain D. Right: Picking points on a circular arc inside a triangle. The line p_1p_3 is tangent to the corresponding circle.

the line that is perpendicular to the line p_1p_3 and passes through p_1. When mirroring p_2 with that line as an axis, the resulting point p_2' is rational and also on Z.) We can now choose any number of rational points on the circular arc by selecting a family of lines through p_1. To this end, we choose a set of equidistant points on the segment p_2p_3, which, together with p_1 define this family of rational lines. Again, the numerators and the denominators are polynomial in those of p_1, p_2, and p_3, and the number of points chosen.

A.2 Constructing Sink Gadgets

We now construct the sink gadgets. See Fig. 11 for an accompanying illustration. Recall that, since β is even, there are no small double chains on neighboring positions on the lower chain. Hence, for each sink we w.l.o.g. can define an orthogonal region within which we can safely draw the small double chain; we call this region the *bin* of the sink (outlined gray in Fig. 11). Consider a sink (i, j). The vertical line bounding the left side of its bin passes through the edge $l_{j-1}l_j$ (e.g., at the midpoint of the edge), and the right side of the bin is defined analogously. (Recall that, since $\beta > 1$, there is no sink at $l_{j-1}l_j$.) Pick a rational point p_a on the boundary of the bin that is to the left of the directed line l_ju_{i-1} and to the right of the directed line l_ju_i. Similarly, choose a point p_b that is to the right of the line $l_{j+1}u_{i+1}$ and to the left of the line $l_{j+1}u_i$. As an additional constraint let p_a be to the left of the line p_bu_i. Note that such points always exist, and can be easily chosen along the boundary of the bin. It remains to choose a triangular region with l_jp_a as one side in which we can place the chain of the sink gadget that contains l_j. For the second chain, the construction is analogous.

For the chain to be visible from u_i but not from u_{i-1}, the triangular region has to be to the left of the line p_au_i, and also to the left of the line l_ju_{i-1}. Further, to be visible from all vertices of the other chain, it has to be to the left of the lines p_al_{j+1} and p_bl_j. Let x_a be the apex of the triangle that is defined by these constraints, and observe that x_a is the intersection of two of the four lines. We can now add a chain of points on a circular arc inside the triangle $l_jp_ax_a$, as described above.

The coordinates are rational, and since every point can be constructed using only a constant number of other points, the numerator and denominator of each point are polynomial.

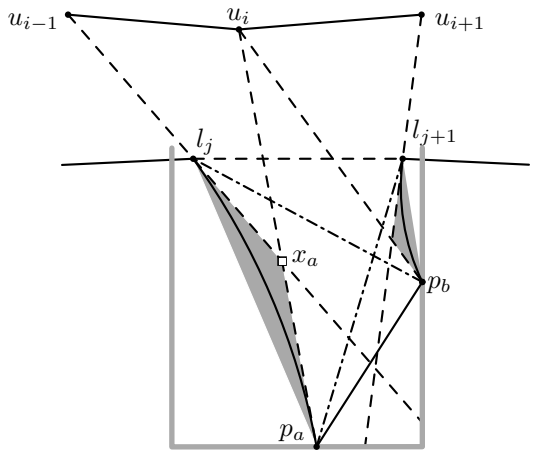

Fig. 11 Construction of a small double chain for a sink.

A.3 General Position and Integer Coordinates

The ways in which a simple polygon can be triangulated is determined by the *order type* of the vertices, i.e., the vector that indicates for each triple of vertices whether it is oriented clockwise or counterclockwise. Up to now, we did not care whether the point set is in general position, so there might also be collinear point triples among vertices that are not directly related in the reduction. By simply multiplying all coordinates by all denominators used, we would obtain integer coordinates with exponential values. To obtain integer coordinates bounded by a polynomial in the input size and a point set in general position, we can use the following lemma.[3]

Lemma A.1 *Let S be a point set with rational coordinates whose numerators and denominators have absolute values of at most ξ. Then there is a point set S' with integer coordinates bounded by $O(\xi^3)$ and a bijection between S and S' such that for every ordered triple of non-collinear points in S, the orientation of the corresponding triple in S' is the same. In particular, if S is in general position, then S and S' have the same order type. Further, S' can be constructed in $O(|S|^2)$ time.*

Proof Consider the set L of lines that are defined by all pairs of points of S. Choose $\ell = q_1q_2 \in L$ and $p \in S \setminus \ell$ such that the horizontal distance v between p and ℓ is minimal among all such distances (which is non-zero as p is not on ℓ). Then v is rational, with numerator and denominator in $O(\xi^2)$. Further, our choice required $v > 0$. When multiplying all x-coordinates by $2/v$, this distance is at least 2. The basic idea is to round the x-coordinates. The crucial observation is that p has a y-coordinate that is between the ones of q_1 and q_2, as otherwise one of q_1 and q_2, say, q_1, would be horizontally closer to the line through p and q_2. For an ordered triple of points to change its orientation (from, say, clockwise to counterclockwise), the horizontal distance between the point whose y-coordinate is between those of the other two points would have to be reduced by more than 2. We can therefore safely round the x-coordinates, which, in the worst case, reduces the horizontal distance between p and ℓ by at most 1. Hence, for every non-collinear ordered triple of points in S, the orientation of the corresponding

[3] The exact time bounds shown in the proof are irrelevant for the NP-hardness reduction (which even requires a different model of computation). We mention them only as they may be of general interest.

triple in the resulting point set is the same. We repeat the process analogously for the y-coordinates, obtaining S'.

The horizontal or vertical distance v can easily be found by checking all triples of points. We can improve this cubic time bound by considering the dual line arrangement $\mathcal{A}$ of S (in which a point $p = (x_p, y_p)$ corresponds to the dual line $p^* : y = x \cdot x_p + y_p$). The dual arrangement can be constructed in quadratic time [6,8]. The shortest vertical distance in the primal corresponds to the shortest vertical distance of a vertex and a side of a triangle defined by three dual lines. Clearly, the shortest distance can only occur inside a triangle that is not intersected by another line.[4] Hence, we only need to test the $O(|S|^2)$ triangular cells of $\mathcal{A}$. □

Hence, if we construct the vertices with rational coordinates such that the vertices are in general position, we can apply Lemma A.1 to have all vertices on the integer grid in general position.

General position can easily be obtained by applying a simple technique used in [19, Appendix A]. Observe first that the vertices of D are in general position. We take special care when placing the $d-1$ points of each chain of a sink gadget to not produce collinear points. Note that the final polygon P will have $|P| = 2(\beta n + 2) + 2N(d-1)$ vertices. Instead of $d-1$ points, we choose $2\binom{|P|}{2} + d - 1$ *candidate points* on the circular arc for the chain. Consider any line through two already placed points. This line intersects the circular arc in at most two points, so there are at most two candidate points that may not be points of the double chain because of that line. As there are less than $\binom{|P|}{2}$ such lines, there are always enough candidate points left for selecting the $d-1$ points for the chain among them. Thus, the vertices we obtain are in general position.

[4] Actually, any dual transform will do. When thinking of the rounding process as a continuous transformation, a change of the order type would involve a collapsing triangular cell of the dual arrangement, indicating a "close" point triple.