

## Algorithms for Routing in Planar Graphs\*

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**Summary.** A routing problem is given by a planar graph  $G=(V, E)$  with a given embedding into the plane and a set  $Ne$  of nets. A net is a pair of points on the boundary of the infinite face. The goal is to find a set of pairwise *edge-disjoint* paths connecting the terminals of the various nets. We assume that the degree of every vertex not on the boundary of the infinite face is even and call such routing problems half-even. We show that one can decide in time  $O(bn)$  whether a half-even problem is solvable and that a solution can be constructed in time  $O(n^2)$ . Here  $n=|V|$  and  $b$  is the number of vertices on the boundary of the infinite face. If the routing problem is even, i.e. every cut has even free capacity, and  $G$  is a subgraph of the planar grid then a solution can be found in time  $O(n^3)$ .

### 1. Introduction

A planar routing problem is given by

- a) a planar undirected graph  $G=(V, E)$  with a fixed embedding into the plane
- b) a set  $Ne$  of two-terminal nets where a net  $N \in Ne$  is an unordered pair of vertices on the boundary of the infinite face.

A solution to a planar routing problem  $(V, E, Ne)$  is given by a set of *pairwise edge-disjoint* paths  $P_N, N \in Ne$ , such that  $P_N$  connects the two terminals of net  $N$ .

For sets  $X, Y \subseteq V$  we define

$$\begin{aligned} \text{cap}(X, Y) &= |\{\{x, y\} \in E : x \in X, y \in Y\}| \\ \text{cap}(X) &= \text{cap}(X, V - X) \\ d(X, Y) &= |\{\{s, t\} \in Ne : s \in X, t \in Y\}| \\ d(X) &= d(X, V - X) \\ \lceil \text{cap}(X) \rceil &= \text{cap}(X) - d(X). \end{aligned}$$

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We call  $\text{fcap}(X)$  the *free capacity* of cut  $X$ . If  $x$  is a vertex we write  $\text{fcap}(x)$ ,  $\text{cap}(x)$ ,  $d(x)$  instead of  $\text{fcap}(\{x\})$ ,  $\text{cap}(\{x\})$  and  $d(\{x\})$ . Note that  $\text{cap}(x)$  is the degree of vertex  $x$ . The following theorem was shown by Okamura et al. [10].

**Theorem** [9]. *Let  $(V, E, Ne)$  be a planar routing problem. If  $\text{fcap}(X)$  is non-negative and even for every  $X \subseteq V$  then the routing problem has a solution.  $\square$*

In this paper we will present algorithms for solving routing problems in planar graphs. We use the following notation.

*Definition.* Let  $P=(V, E, Ne)$  be a planar routing problem,  $n=|V|$ . Let  $b$  be the number of vertices on the boundary of the infinite face. The problem  $P$  is *solvable* if it has a solution. It is *even* if  $\text{fcap}(x)$  is even for all  $x \in V$ . It is *half-even* if  $\text{fcap}(x)=\text{cap}(x)$  is even for all nodes  $x$  which do not lie on the boundary of the infinite face.  $\square$

We prove in this paper the following

**Theorem.** *Let  $P=(V, E, Ne)$  be a half-even routing problem.*

- a) *It can be tested in time  $O(bn)$  whether  $P$  is solvable.*
- b) *A solution can be constructed in time  $O(n^2)$ .*
- c) *If  $(V, E)$  is a grid graph, i.e. a subgraph of the planar integer grid, then a solution can be found in time  $O(bn)$ . If  $P$  is even then time  $O(n^{3/2})$  suffices.*

A proof of parts a), b) and c) for even problems can be found in Sect. 2. The extension to half-even problems is made in Sect. 3. In Sect. 4, a weak generalization of the theorems above to multi-terminal nets is made. If  $2d(X) < \text{cap}(X)$  for all cuts  $X$  then a solution exists and can be found in time  $O(n^2)$ .

We close the introduction with a comparison to other work. Hassim [3] and Matsumoto et al. [6] consider the problem of multi-commodity flows in planar networks, i.e. edges have real capacities and nets have real demands. The goal is to construct flow functions which realize the demands and obey the capacity constraints. Our problem differs from the multicommodity flow problem in two respects

a) it is simpler since all capacities and demands are equal to one in our case.

b) it is harder since we insist that a solution consists of edge-disjoint paths. Figure 1 shows a problem which has a solution as a flow problem (all edges carry flow 1/2 for both commodities) but not as a routing problem.

Hassim [3] and Matsumoto et al. [6] describe algorithms which solve multi-commodity flow problems in planar graphs in time  $O(n^4)$  and time  $O(kn + n^2 \sqrt{\log n})$  respectively where  $k=|Ne|$ . If all capacities and demands are integral and the problem is even then the flow functions constructed by their algorithms are integral. In particular, even routing problems can be solved in time  $O(n^2 \sqrt{\log n})$  by their algorithms; note that  $k \leq |E| = O(n)$  in planar routing problems. Matsumoto et al. [6] also show how to test solvability in time  $O(\min(n^2 \log^* n, bn(\log n)^{1/2}))$ . Our improvement to their results for even prob-

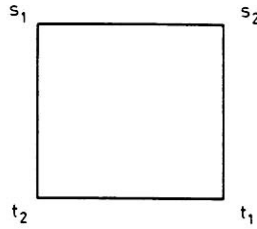


Fig. 1. A problem which is solvable as a flow problem but not as a routing problem: the two nets are  $(s_1, t_1)$  and  $(s_2, t_2)$

lems is trivial; we only show how to drop the various factors involving  $\log n$ . However, our algorithm for even problems differs considerably from theirs. It has the advantage that it can be made to run in time  $O(n^{3/2})$  on grid graphs while their algorithm will always run in time  $O(n^2)$ . The other major contribution of the present paper is the extension of the theory to half-even problems. This extension was obtained independently by Frank [2]. However, the running time of his algorithm is  $O(n^3)$  whilst the running time for our algorithm is  $O(bn) = O(n^2)$ . The problem of extending our results beyond half-even problems remains as a major challenge.

Grid graphs are particularly relevant for VLSI routing problems. Special cases of grid graph routing problems were considered previously by Rivest et al. [12], Preparata et al. [11], Frank [1], Mehlhorn et al. [8], Kaufmann et al. [4] and Nishizeki/Saito/Suzuki. They show how to solve problems for channels, rectangles, grid graphs without holes, and convex grid graphs without holes.

Figure 6 shows a grid graph which has an interior face with more than four vertices and hence cannot be handled by the last two papers mentioned.

## II. Even Routing Problems

Throughout this section  $P=(V, E, Ne)$  is an even routing problem. We will show how to construct a solution (if there is one) in time  $O(n^2)$ . For grid graphs time  $O(n^{3/2})$  suffices. Our algorithm is an almost direct implementation of Okamura et al. [10] constructive proof of their theorem and hence correctness can be shown using their methods.

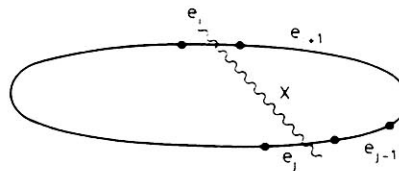


Fig. 2. A cut  $X$  through edges  $e_i$  and  $e_{i+1}$

**Theorem 1.** Let  $P=(V, E, Ne)$  be an even routing problem. Then it can be tested in time  $O(bn)$  whether  $P$  has a solution. If  $P$  has a solution then a solution can be constructed in time  $O(n^2)$ .

*Proof.* We will first show how to construct a solution in time  $O(n^2)$ . The test for solvability will be a corollary.

We need one additional definition. A cut  $X$  is *simple* if there are exactly two edges  $e', e''$  on the boundary of the infinite face having exactly one endpoint in  $X$  (cf. Fig. 2). We are now ready for the algorithm.

- (1) **while**  $E \neq \emptyset$
  - (2) **do** let  $e=(a, b)$  be an edge on the boundary of the infinite face:
  - (3) let  $e_0, e_1, \dots, e_{m-1}$  be the edges on the boundary of the infinite face in clockwise order with  $e=e_0$ ;
  - (4) **if** there is a simple cut  $X$  with  $a \notin X, b \in X$  with  $f \text{ cap}(X) < 0$  **then** halt and declare the problem unsolvable **fi**;
  - (5) **if** there is a simple cut  $X$  with  $a \notin X, b \in X$  and  $f \text{ cap}(X) = 0$
  - (6) **then** let  $X$  be such a simple cut which contains as few vertices of the boundary as possible:  
 let  $N=(s, t)$  be a net with  $s \notin X, t \in X$  and  $s$  as close to  $a$  as possible, i.e. there is no net  $(s', t')$  with  $s' \notin X, t' \in X$  and  $s'$  lying on that piece of the boundary between  $s$  (exclusive) and  $a$  (inclusive) not containing  $b$  (cf. Fig. 3); delete edge  $e$ , reserve edge  $e$  for net  $N$  and replace net  $N$  by nets  $N_1=(s, a)$  and  $N_2=(b, t)$ .  
**else** delete edge  $e$  and add an additional net  $N=(a, b)$ .  
**fi**
- od**

Figure 10 shows a sample execution of the algorithm.

The correctness of this algorithm is almost immediate from the work of Okamura et al. [10]. We give a short correctness proof to make the paper self-contained.

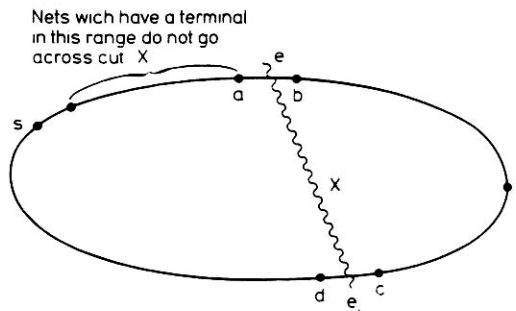


Fig. 3.  $f \text{ cap}(e, e) = f \text{ cap}(X) = 0$

Let us assume that  $P$  is solvable. Then clearly  $f\text{cap}(X) \in 2\mathbb{N}_0$  ( $2\mathbb{N}_0$  denotes the even non-negative integers) for every simple cut  $X$  initially. It suffices to show that  $f\text{cap}(X) \in 2\mathbb{N}_0$  for every simple cut  $X$  is an invariant of the algorithm because then line (4) is never executed and the algorithm will construct a solution.

Consider an execution of the while-loop which removes edge  $e$  from the current graph  $G_c$ . Note first  $f\text{cap}(Y)$  is even for all cuts  $Y$  iff  $f\text{cap}(v)$  is even for every vertex  $v$ . This can be seen as follows. Let  $Y \subseteq V$  be arbitrary. Then

$$\begin{aligned} f\text{cap}(Y) &= \text{cap}(Y) - d(Y) \\ &= \sum_{v \in Y} \text{cap}(v) - 2 \text{cap}(Y, Y) - \sum_{v \in Y} d(v) + 2d(Y, Y) \end{aligned}$$

and hence  $f\text{cap}(Y)$  is even if  $f\text{cap}(v)$  is even for all  $v \in V$ . It is now easy to see that all problems constructed by the algorithm are even.

Let  $Y$  be a simple cut in  $G_c - e$ . Then  $Y$  is also a simple cut in  $G_c$  and  $f\text{cap}(Y) \geq \overline{f\text{cap}}(Y) - 2$  where  $\overline{f\text{cap}}$  is the free capacity in the modified problem. If  $f\text{cap}(Y) \geq 2$  or  $\overline{f\text{cap}}(Y) = f\text{cap}(Y)$  then clearly  $\overline{f\text{cap}}(Y) \in 2\mathbb{N}_0$ . Recall that we have already shown that the modified problem is even. The only case to consider is therefore that  $f\text{cap}(Y) = 0$  and  $\overline{f\text{cap}}(Y) = f\text{cap}(Y) - 2$ . We may assume w.l.o.g. that  $t \notin Y$ ; consider  $V - Y$  otherwise. If  $\overline{f\text{cap}}(Y) = f\text{cap}(Y) - 2$  then either (Case A)  $|Y \cap \{a, b\}| = 1$  and  $s \notin Y$  or (Case B)  $a, b \in Y$  and  $s \notin Y$ . We can restrict Case A further. If  $a \notin Y, b \in Y$  then  $Y$  contains fewer boundary nodes than  $X$ , a contradiction to the choice of  $X$ . We may therefore assume  $a \in Y, b \notin Y$  in Case A. The following lemma was shown in Okamura et al. [10].

**Lemma 1.** For all cuts  $X$  and  $Y$

$$\begin{aligned} d(X) + d(Y) &= d(X \cup Y) + d(X \cap Y) + 2d(X - Y, Y - X) \\ \text{cap}(X) + \text{cap}(Y) &= \text{cap}(X \cup Y) + \text{cap}(X \cap Y) + 2 \text{cap}(X - Y, Y - X). \quad \square \end{aligned}$$

We will apply Lemma 1 in both cases. The following observation is also used in both cases. The boundary nodes in  $Y - X$  lie between  $s$  (exclusive) and  $a$  (inclusive) and hence  $d(X - Y, Y - X) = 0$  by choice of net  $N = (s, t)$ . We consider cuts  $X \cup Y$  and  $X \cap Y$ . We have

$$\begin{aligned} d(X \cup Y) + d(X \cap Y) &= d(X) + d(Y), && \text{by Lemma 1 and the} \\ & && \text{observation above} \\ &= \text{cap}(X) + \text{cap}(Y), && \text{since } X \text{ and } Y \text{ are} \\ & && \text{saturated} \\ &= \text{cap}(X \cup Y) + \text{cap}(X \cap Y) + 2 \text{cap}(X - Y, Y - X), && \\ & && \text{by Lemma 1.} \end{aligned}$$

*Case A.* In Case A we have  $a \in Y - X, b \in X - Y$  and hence  $\text{cap}(X - Y, Y - X) > 0$  and thus  $d(X \cup Y) + d(X \cap Y) > \text{cap}(X \cup Y) + \text{cap}(X \cap Y)$ . Thus one of the simple cuts  $X \cup Y$  or  $X \cap Y$  is oversaturated, a contradiction.

*Case B.* In Case B we conclude  $d(X \cup Y) + d(X \cap Y) \geq \text{cap}(X \cup Y) + \text{cap}(X \cap Y)$ . Thus  $f \text{cap}(X \cup Y) = 0 = f \text{cap}(X \cap Y)$ . Next note that  $b \in X \cap Y$  and that simple cut  $X \cap Y$  contains fewer boundary nodes than  $X$ . This contradicts the choice of  $X$ .

This completes the correctness proof. Note that the correctness proof also shows that  $P$  is solvable iff  $f \text{cap}(X) \geq 0$  for every simple cut in the initial problem. This observation will lead to the efficient test for solvability.

We turn to the implementation next. The main task is to determine the existence of a saturated cut through edge  $e = (a, b)$ . As in Hassim [3] and Matsumoto et al. [6] we solve this task by means of the multiple source dual graph (cf. Fig. 4). In the dual graph there is a dual edge for every edge of the original graph. The dual edge connects vertices which are located in the faces separated by the edge. In every face (except the infinite face) we position one dual vertex but in the infinite face we have a dual vertex for every edge on the boundary of the infinite face.

Let  $e_0, \dots, e_{m-1}$  be the edges on the boundary of the infinite face in clockwise order and let  $e = e_0$ . Let

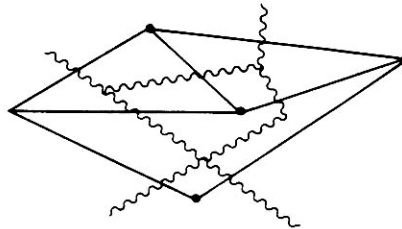
$$\begin{aligned} \text{cap}(e, e_i) &= \min \{ \text{cap}(X); X \text{ is a simple cut which cuts} \\ &\quad \text{boundary edges } e \text{ and } e_i \} \\ d(e, e_i) &= \{ d(X); X \text{ is a simple cut which cuts} \\ &\quad \text{boundary edges } e \text{ and } e_i \}, \text{ and} \\ f \text{cap}(e, e_i) &= \text{cap}(e, e_i) - d(e, e_i). \end{aligned}$$

Let  $v_i$  be the dual vertex in the infinite face corresponding to edge  $e_i$ . Then  $\text{cap}(e, e_i)$  is equal to the length of a shortest path from  $v_0$  to  $e_i$  in the dual graph and hence  $\text{cap}(e, e_i), 1 \leq i \leq m-1$ , can be computed in time  $O(n)$  by breadth first search. It is also easy to see that  $d(e, e_i), 1 \leq i \leq m-1$ , can be computed in time  $O(n)$  by a simple walk around the boundary of the infinite face. We summarize in

**Lemma 2.**  $f \text{cap}(e, e_i), 1 \leq i \leq m-1$ , can be computed in time  $O(n)$ .  $\square$

The remainder of the loop body can clearly also be done in time  $O(n)$ . Thus a single execution of the loop body takes time  $O(n)$  and hence total running time is  $O(n^2)$ .

For the test of solvability we only have to compute  $f \text{cap}(e_i, e_j)$  for all  $i$  and  $j$ . For every fixed  $i$  this takes time  $O(n)$ . Thus total running time is  $O(hn)$  where



**Fig. 4.** A planar graph  $G$  and its multiple-source dual. Dual edges are shown as wiggled lines

$b$  is the number of edges on the boundary of the infinite face in the initial graph.  $\square$

We will next show how to improve upon Theorem 1 for planar graphs with small  $C$ -connected edge separators.

*Definition.* a) Let  $G=(V, E)$  be a planar graph. A set  $E' \subseteq E$  of edges is a  $C$ -connected edge separator if

1) removal of  $E'$  splits  $G$  into subgraphs  $(V_1, E_1)$  and  $(V_2, E_2)$  with  $|V_1| \leq 2.3|V|$  and  $|V_2| \leq 2.3|V|$

2)  $E'$  can be ordered, say  $E'=(e_1, \dots, e_k)$  such that edge  $e_i$  is on the boundary of the infinite face of graph  $G - \{e_1, \dots, e_{i-1}\}$ .

b) A family  $F$  of planar graphs has  $C$ -connected edge separators of size  $f(n)$  if every  $G=(V, E) \in F$  has a  $C$ -connected edge separator  $E'$  of size  $|E'| \leq f(|V|)$  and if the subgraphs  $(V_1, E_1)$  and  $(V_2, E_2)$  obtained by removing  $E'$  again belong to  $F$ .  $\square$

$C$ -connected edge separators are a very restrictive notion of separator which is particularly suitable for our application. However, not all planar graphs have small  $C$ -connected edge separators as Fig. 5 shows. The graph shown in Fig. 5 consists of  $n$  nested triangles. All  $C$ -connected edge separators have size  $\Omega(n)$ . Fortunately, there is a family of graphs having small  $C$ -connected separators, namely grid graphs.

*Definition.* A *grid graph* is a subgraph of the two-dimensional integer grid.  $\square$

Figure 6 shows a grid graph.

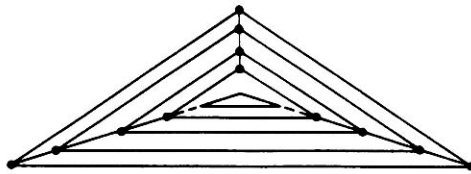


Fig. 5. A planar graph all of whose  $C$ -connected edge separators have size  $\Omega(n)$

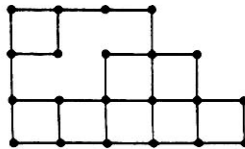


Fig. 6. A grid graph

**Lemma 3.** Let  $G$  be a grid graph with  $n$  vertices. Then  $G$  has a  $C$ -connected edge separator of size  $\leq 4\sqrt{n+2}$ . Moreover, the separator can be found in time  $O(n)$ .

*Proof* (adapted from Lengauer et al. [5], Theorem 1). Let integers  $a$  and  $b$  be minimal such that a rectangle of side lengths  $a$  and  $b$  encloses the grid graph.

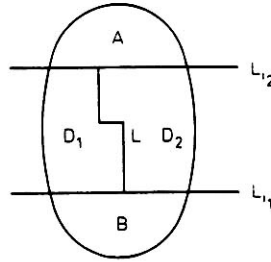


Fig. 7. Cuts  $L_{i_1}$ ,  $L_{i_2}$  and  $L$

Assume w.l.o.g. that  $a \leq b$  and that the side of length  $a$  is horizontal. Let  $L_i$ ,  $-1 \leq i \leq b$ , be the number of edges intersected by a horizontal line which runs in distance  $i + 1/2$  from the bottom side of the rectangle. Then  $L_{-1} = L_b = 0$  and  $\sum_i L_i \leq n$ . Let  $i_0$  be minimal such that at least  $n/2$  nodes lie below line  $L_{i_0}$ , i.e.  $\leq n/2$  nodes lie below line  $L_{i_0-1}$  and  $\leq n/2$  nodes lie above line  $L_{i_0}$ . Let  $i_1 < i_0$  and  $i_2 \geq i_0$  be such that (cf. Fig. 7)

- 1)  $i_1 \geq i_0 - 1 - \sqrt{n}$  and  $L_{i_1} \leq \sqrt{n}$ .
- 2)  $i_2 \leq i_0 + \sqrt{n}$  and  $L_{i_2} \leq \sqrt{n}$ .

Note that  $i_1$  exists because  $\sum_{i_0-1-\sqrt{n} \leq i \leq i_0-1} L_i \leq n$ ;  $i_2$  exists for a similar reason.

Let  $A$  be the set of vertices above line  $L_{i_2}$ ,  $B$  be the set of vertices below line  $L_{i_1}$  and  $C$  be the set of vertices between the two lines. Let  $L$  be a vertical line with maybe one horizontal segment of length one running between  $L_{i_1}$  and  $L_{i_2}$  and dividing  $C$  into two parts  $C_1$  and  $C_2$  of size  $\lfloor |C|/2 \rfloor$  and  $\lceil |C|/2 \rceil$  respectively. Then  $|A| \leq \lceil n/2 \rceil$ ,  $|B| \leq \lceil n/2 \rceil$ ,  $|C_1| \leq \lceil n/2 \rceil$  and  $|C_2| \leq \lceil n/2 \rceil$ . Let  $V_1$  be the largest of the four sets  $C_1, C_2, A, B$  if this set has size  $\geq 1/3 n$  and let  $V_1$  be the union of the two largest sets otherwise. Let  $V_2 = V - V_1$ . Then  $|V_1| \leq 2n/3$ ,  $|V_2| \leq 2n/3$  and hence the set of edges intersected by lines  $L_{i_1}, L_{i_2}$  and  $L$  forms a  $C$ -connected edge separator. This set has size  $\leq 4\sqrt{n} + 2$ .  $\square$

**Theorem 2.** Let  $P = (V, E, Ne)$  be an even routing problem with  $(V, E)$  a grid graph. Then a solution can be constructed (if there is one) in time  $O(n^{3/2})$ .

*Proof.* Let  $E' \subseteq E$  be a  $C$ -connected edge separator of size  $|E'| = O(\sqrt{n})$ ;  $E'$  exists by Lemma 3. Let  $E' = (e_1, \dots, e_k)$  where  $e_i$  is on the boundary of the infinite face of graph  $G - \{e_1, \dots, e_{i-1}\}$ . We use the algorithm described in the proof of Theorem 1 and let  $e$  run through edges  $e_1, e_2, \dots, e_k$  in the first  $k = O(1)$  iterations of the loop. This takes time  $O(n^{3/2})$  and splits the graph into subgraphs with  $n_1$  and  $n_2$  vertices respectively, where  $\max(n_1, n_2) \leq 2n/3$ . We therefore have the following recurrence for the running time

$$T(n) \leq O(n^{3/2}) + \max \{ T(n_1) + T(n_2); n_1 + n_2 = n, \max(n_1, n_2) \leq 2n/3 \}.$$

Thus  $T(n) = O(n^{3/2})$  as claimed (cf. Mehlhorn [7], p. 121 where a similar recurrence is solved).  $\square$



### III. Half-even Problems

A routing problem  $P=(V, E, Ne)$  is *half-even* if  $f \text{cap}(v)=\text{cap}(v)=\text{deg}(v)$  is even for all nodes not on the boundary of the infinite face. Figure 1 shows a half-even routing problem. In this section we extend our results to half-even routing problems. We proceed in three steps. In the first step (Lemma 4a) we show that a solvable half-even problem  $P=(V, E, Ne)$  always has a solvable even extension  $P=(V, E, Ne \cup Ne')$ , in the second step (Lemma 4b), we develop one particular strategy for computing such an extension and in the third step (Theorem 3) we show how to make the strategy run in time  $O(bn)$ . A similar but less efficient technique was developed independently by Frank [2].

Our reduction is based on the concepts of *U-minimal cut* and *canonical extension*. Let  $U=\{v: f \text{cap}(v) \text{ is odd}\}$  be the vertices with odd free capacity. Then  $U$  has even cardinality and all vertices in  $U$  lie on the boundary of the infinite face (recall that we deal with half-even problems). Let  $X$  be a saturated cut. Then  $X \cap U$  has even cardinality (this can be seen as follows:  $O = \text{cap}(X) - d(X) = \sum_{x \in X} (\text{cap}(x) - d(x)) - 2 \text{cap}(X, \bar{X}) + 2d(X, \bar{X})$ . Thus the number of vertices in  $X$  with odd free capacity is even). Let  $u_1, \dots, u_{2k}$  be the vertices in  $X \cap U$  in clockwise order on the boundary of the infinite face. Cut  $X$  is *U-minimal* if  $X \cap U \neq \emptyset$  and there is no simple saturated cut  $Y$  with  $Y \cap U = \{u_i, \dots, u_j\}$  where  $1 < i < j < 2k$ . The *canonical extension* of  $P=(V, E, Ne)$  with respect to  $X$  is given by  $P'=(V, E, Ne \cup \{(u_{2i-1}, u_{2i}): 1 \leq i \leq k\})$ . Note that all vertices of  $X$  have even free capacity in  $P'$ .

**Lemma 4.** *Let  $P=(V, E, Ne)$  be a solvable half-even routing problem.*

- a) *There is a solvable even extension  $(V, E, Ne \cup Ne')$  where  $Ne'$  is a pairing of  $U = \{v: f \text{cap}(v) \text{ is odd}\}$*
- b) *If  $X$  is a U-minimal cut then the canonical extension of  $P$  with respect to  $X$  is a half-even solvable routing problem.*

*Proof.* a) Let  $P_N, N \in Ne$ , be a solution for  $P$ . Consider graph  $G'=(V, E')$  with  $E' = E - \{P_N; N \in Ne\}$  which is obtained from  $G=(V, E)$  by removing all edges which are used in the solution paths. We have

- a) if  $v$  is a vertex on the boundary of  $G$  then  $v$  has odd degree in  $G'$  iff  $v \in U$
- b) if  $v$  is a vertex in the interior of  $G$  then  $v$  has even degree in  $G'$ .

We conclude that  $G'$  decomposes into paths connecting vertices in  $U$  and cycles. The paths connecting vertices in  $U$  induce the desired pairing  $Ne'$ . This proves a).

b) Let  $P_0=(V, E, Ne \cup Pa)$  be a solvable even extension of the half-even problem  $P=(V, E, Ne)$  and let  $X$  be a  $U$ -minimal cut. We have argued above that  $U \cap X$  is even. Let

$$Pa' = \{(s, t) \in Pa: s \in X, t \notin X\}.$$

Let  $P'=(V, E, Ne \cup Pa')$ . Then  $P'$  is a half-even solvable problem with  $U \cap X$  as its set of vertices of odd free capacity. Let  $P''$  be the canonical extension of  $P'$

with respect to  $X$ . Then  $P''$  is an even problem which extends the canonical extension of  $P$  with respect to  $X$ . It therefore suffices to show that  $P''$  is solvable.

Let  $Y$  be a simple cut. Then  $d''(Y) \neq d'(Y) + 2$  where  $d''(Y)$  ( $d'(Y)$ ) denotes the density of  $Y$  with respect to problem  $P''$  ( $P'$ ). Also  $d'(Y) \leq \text{cap}'(Y)$  since  $P'$  is solvable. Assume first that  $d''(Y) \neq d'(Y) + 1$ . Then  $\text{cap}''(Y) - d''(Y) \geq \text{cap}'(Y) - d'(Y) - 1 \geq -1$ . Since  $P''$  is even conclude  $\text{cap}''(Y) - d''(Y) \geq 0$ . This leaves the case that  $d''(Y) = d'(Y) + 2$ . Then  $Y \cap U = \{u_i, \dots, u_j\}$  for some  $i, j$  with  $1 < i < j < 2k$  where  $u_1, \dots, u_{2k}$  is the clockwise ordering of the vertices in  $U \cap X$ . Thus  $Y$  is not saturated in  $P$  since  $X$  is a  $U$ -minimal cut. Also the nets in  $Pa'$  pair only vertices outside  $Y$ . Thus  $Y$  is not saturated in problem  $P'$  and therefore  $\text{cap}''(Y) - d''(Y) \geq \text{cap}'(Y) - d'(Y) - 2 \geq 1 - 2 = -1$ . Since  $P''$  is even this implies  $\text{cap}''(Y) - d''(Y) \geq 0$ .

We conclude that there are no oversaturated simple cuts in  $P''$ . Hence  $P''$  is solvable.  $\square$

Lemma 4b leads to the following algorithm for turning half-even problems into even problems.

- (1)  $U \leftarrow \{v; f \text{ cap}(v) \text{ is odd}\}$
- (2) **while**  $U \neq \emptyset$
- (3) **do if** there is an oversaturated cut
- (4) **then** terminate and declare that the problem has no integer-valued solution
- (5) **fi**:
- (6) let  $X$  be a  $U$ -minimal cut ( $X = V$  is possible);
- (7) construct the canonical extension:
- (8)  $U \leftarrow U - X$
- (9) **od**

The correctness of this algorithm is immediate from Lemma 4b. Note that  $U \cap X$  has always even cardinality and hence line (7) is always executable. Figure 8 illustrates the extension algorithm.

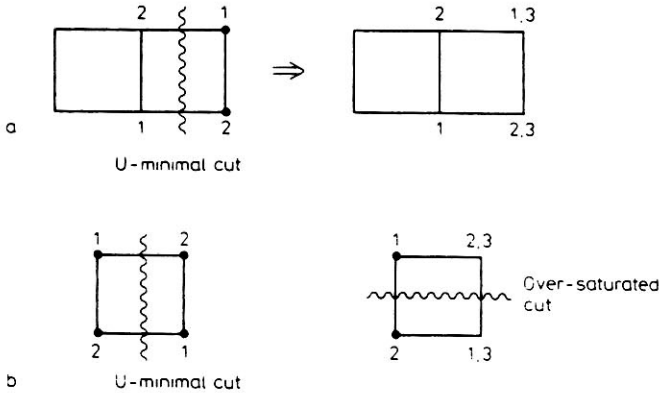


Fig. 8a and b. The extension algorithm: odd nodes are shown solid. a A solvable problem, b an unsolvable problem

It remains to discuss an efficient implementation. The most difficult design decision is how to handle  $U$ -minimum cuts. We start with the observation that if  $X$  and  $X'$  are cuts with  $U \cap X = U \cap X'$  and  $f \text{cap}(X) \leq f \text{cap}(X')$  then only cut  $X$  has to be considered because of the following two trivial facts.:

1) if net  $(s, t)$  is added in line (7) then  $f \text{cap}(X)$  goes down by one iff  $f \text{cap}(X')$  does.

2) if  $X'$  is  $U$ -minimal then  $X$  is  $U$ -minimal.

The observation above suggests the following strategy for finding  $U$ -minimal cuts. Let the vertices on the boundary of the exterior face be labelled  $0, 1, \dots, l-1$  in clockwise order and let  $U = \{u_0, u_1, \dots, u_{2k-1}\}$ . For  $0 \leq p, q \leq 2k-1$  let  $X(p, q)$  be a simple cut with

1)  $X(p, q) \cap U = \{u_p, u_{p+1}, \dots, u_q\}$ .

2)  $X(p, q)$  has the smallest free capacity among all cuts satisfying 1). Ties are broken arbitrarily.

Note that we may assume w.l.o.g. that  $X(p, q) = X(q+1, p-1)$ . We represent cut  $X(p, q)$  by the triple  $(p, q, f)$  where  $f = f \text{cap}(X(p, q))$ . We store the cuts  $X(p, q)$  as follows.

For every  $p$  have the linked list of (representatives of) cuts  $X(p, q)$  in clockwise order of  $q$ .

In addition, we have a pointer  $\text{min}(p)$  pointing to the first  $q$  such that  $X(p, q)$  is saturated. Furthermore, we link the two occurrences of a cut (namely as  $X(p, q)$  and as  $X(q+1, p-1)$ ). We have

**Lemma 5.** a) The family  $X(p, q)$  of cuts can be constructed in time  $O(bn)$

b) A  $U$ -minimum cut  $X$  can be found in time  $O(|U|)$  given the data structure described above

c) The data structure can be updated in time  $O(|U|)$  after adding net  $(u_i, u_{i-1})$  for some  $i$ .

*Proof.* a) Immediate from the proof of Theorem 1.

b) If there is no  $p$  such that  $\text{min}(p)$  is defined then  $X=V$  is the only saturated cut and hence the only  $U$ -minimum cut. Assume now that  $\text{min}(p)$  is defined for some  $p$ . say  $p=p_0$ . Then the following algorithm finds a  $U$ -minimal cut in time  $O(|U|)$ .

```

p ← p0 + 1    indices are mod 2K
while p lies between p0 and min(p0)
do if min(p) is defined and min(p) lies between p and
    min(p) in the clockwise ordering
    then p0 ← min(p)
fi:
p ← p + 1
od
    
```

c) Suppose that we add net  $(u_i, u_{i-1})$  for some  $i$  and hence delete vertices  $u_i$  and  $u_{i-1}$  from  $U$ . Three actions are required.

1) Reduce  $f\text{cap}(i+1, q)$  and  $f\text{cap}(q, i)$  by one for all  $q$ . Update the min-pointers  $\min(q)$  for all  $q$ . This takes time  $O(|U|)$ .

2) Among the cuts  $X(i, q)$ ,  $X(i+1, q)$  and  $X(i+2, q)$  select the one with smallest free capacity for all  $q$ . This defines the new linked list for vertex  $u_{i+2}$  and takes time  $O(|U|)$ .

3) For every  $q$  keep only one of the cuts  $X(q, i-1)$ ,  $X(q, i)$ , and  $X(q, i+1)$ . This reduces the length of the list for vertex  $u_q$  by 2 and takes time  $O(|U|)$ .

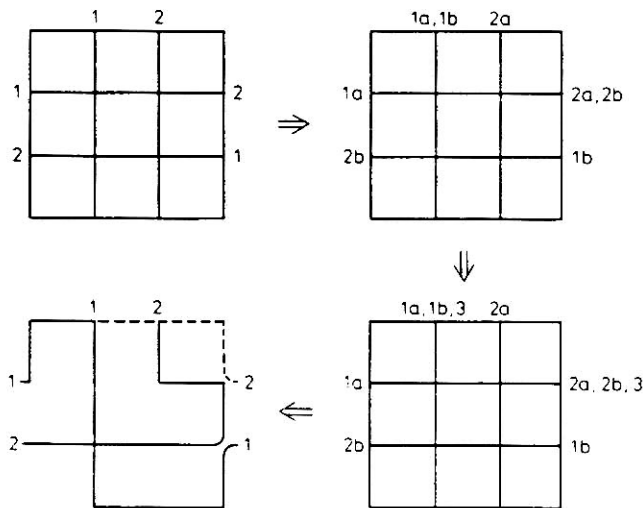
**Theorem 3.** *Let  $P=(V, E, Ne)$  be a half-even routing problem. Then in time  $O(bn)$  one can*

- a) *decide whether  $P$  is solvable and*
- b) *extend  $P$  to a solvable even routing problem  $P'$  if  $P$  is solvable.*

*Proof.* By the algorithm and Lemma 5 above. Note that  $|U| \leq b$  and that the while-loop is executed at most  $b$  times.  $\square$

#### IV. Multi-terminal Nets

Let  $G=(V, E)$  be a planar graph. A multi-terminal net  $N$  is a subset (of size  $\geq 2$ ) of the vertices of the boundary of the infinite face. A routing for a multi-terminal net is a tree  $T(N) \subseteq E$  connecting the points in  $N$ . A routing problem for multi-terminal nets is given by a planar graph  $G=(V, E)$  and a set  $Ne = \{N_1, \dots, N_M\}$  of multi-terminal nets. A solution is given by routings  $T(N_1), \dots, T(N_M)$  such that  $T(N_i)$  is a routing for  $N_i$  and  $T(N_i)$  and  $T(N_j)$  are edge-disjoint for  $i \neq j$ .



**Fig. 9.** A routing problem with multi-terminal nets, a related problem with two-terminal nets, the even problem obtained from Theorem 6, and solution obtained by Theorem 4 (the path for the artificial net 3 is shown dashed)

The density of a cut  $X$  is defined in complete analogy to the case of two-terminal nets, i.e.  $d(X)$  is the number of nets having one but not all terminals in  $X$

$$d(X) = |\{N \in Ne : \emptyset \neq X \cap N \neq N\}|.$$

As before, we call a problem  $(V, E, Ne)$  half-even if  $cap(v) = |\{w \in V : (v, w) \in E\}|$  is even for all interior nodes  $v$ . We have:

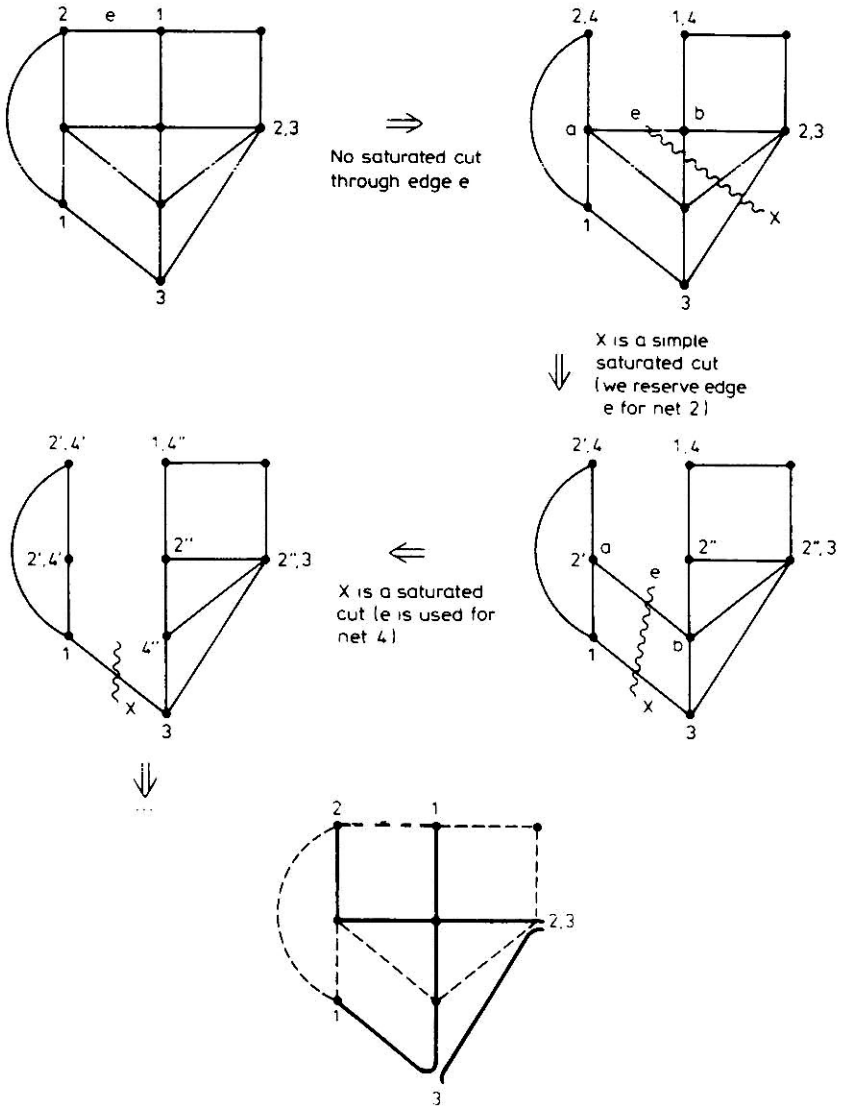


Fig. 10. A sample execution of the algorithm and the solution constructed. The dashed edges realize the net 4 which was introduced in the first iteration

**Theorem 4.** Let  $P=(V, E, Ne)$  be a half-even routing problem with multi-terminal nets. If

$$2 \text{ dens}(X) < \text{cap}(X)$$

for every cut  $X$  then  $P$  has a solution. Moreover a solution can be found in time  $O(n^2)$ . For grid graphs time  $O(n^{3/2})$  suffices.

*Proof.* Let  $Ne = \{N_1, \dots, N_M\}$ . Consider a net  $N_i$ . Let  $v_1, v_2, \dots, v_k$  be the terminals of  $N_i$  as they appear in clockwise order on the boundary of the infinite face. Replace  $N_i$  by the set of  $k-1$  two-terminal nets  $N_i^j = (v_j, v_{j+1})$ ;  $1 \leq j < k$ . Let  $Ne'$  be the new set of nets and let  $\text{dens}'(X)$  be the density of a cut  $X$  with respect to family  $Ne'$  of nets. Then  $\text{dens}'(X) \leq 2 \text{ dens}(X) < \text{cap}(X)$ . The problem  $(V, E, Ne')$  is half-even and has no saturated non-trivial cut. Therefore,  $X=V$  is the only  $U$ -minimal cut and hence  $(V, E, Ne')$  can be turned into an even problem in time  $O(n)$  by the algorithm of Sect. III. The algorithms of Sect. II can then be used to solve the even problem.  $\square$

Figure 9 illustrates the proof of Theorem 4.

## V. Conclusion

We presented efficient algorithms for routing problems in planar graphs. The algorithms are guaranteed to find a solution if there is one. A weak generalization to multi-terminal nets was made. Multi-terminal nets deserve further investigation.

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